Embedding the Pi-Calculus into a Concurrent Functional Programming Language *

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Abstract. We investigate translations from the synchronous pi-calculus into a core language of Concurrent Haskell (CH). Synchronous message-passing of the pi-calculus is encoded as sending messages and adding synchronization using Concurrent Haskell’s mutable shared-memory locations (MVars). Our correctness criterion for translations is invariance of may- and should-convergence. This embraces that all executions of a process are error-free if and only if this also holds for the translated program. We exhibit a particular correct translation that uses a fresh, private MVar per communication interaction and that is in addition adequate, and which is also fully abstract on closed expressions. A meta-result is that CH has the expressive power and the concurrency capabilities of the synchronous pi-calculus.

We also automatically check variants of translations of synchronous communication into an asynchronous calculus where only an a priori fixed number of MVars per channel (and not per communication interaction!) is available. We obtain non-correctness results for classes of small translations, and exemplary argue for the correctness (and adequacy) for two translations with a higher number of MVars. We introduce a classification of the potentially correct translations.

Keywords: pi-calculus, functional programming, concurrency, adequate translations

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1 Introduction

Motivation and Goals. We are interested in the semantics of concurrent and declarative programming languages, where we focus on the expressiveness of the languages and their primitives. A powerful method is the investigation of translations between the programming languages and calculi, which indicate their relative expressiveness. We will mainly compare the $\pi$-calculus with Concurrent Haskell [16,8].

The well-known $\pi$-calculus [13,12,26] is a minimal model for mobile and concurrent processes. Data-flow is possible by communication between processes, i.e. by passing messages between them. Channel names are sent as messages, and processes and links between processes can be dynamically created and removed which together makes processes mobile. The interest in the $\pi$-calculus is not only due to the fact that it is used and extended for various applications, like reasoning about cryptographic protocols [1], applications in molecular biology [18], and distributed computing [10,6]. The $\pi$-calculus also permits the study of intrinsic principles and semantics of concurrency, of concurrent programming and the inherent nondeterministic behavior of mobile and communicating processes. The investigated variant of the $\pi$-calculus in this paper is the synchronous $\pi$-calculus with replication, but without sums, matching operators, or recursion. To observe termination of a process, the investigated $\pi$-calculus has a constant Stop which signals successful termination.

We motivate the use of Concurrent Haskell as a target language, and why usual sequential Haskell (or even Haskell extended by erratic non-determinism) is insufficient: a reason is that the degree of expressiveness for concurrency can be assessed by so-called parallel-convergence testing [19,17]. In the $\pi$-calculus with Stop the context $(\cdot|\cdot)$ is a parallel-convergence tester. Deterministic languages like Haskell do not have parallel-convergence testers. Core-Haskell with erratic choice (i.e., a construct that permits to choose between two expressions) appears to be more expressive, however, this and similar calculi do not have a parallel-convergence testing context $P[\cdot,\cdot]$. Informally, the reason is that if $P[v,\bot]$ as well as $P[\bot,v]$ converge for values $v$, then in such programming languages $P[e_1,e_2]$ has to converge independently of expressions $e_1,e_2$ (in particular, $P[\bot,\bot]$ must also converge). There are stronger nondeterministic primitives, like McCarthy’s amb-operator [11], which can express parallel-convergence testing contexts (see e.g. [21]). Concurrent Haskell [16,8] is very powerful and has a parallel-convergence testing context.

For our technical investigation, we will introduce $CH$ that is a core calculus modeling Concurrent Haskell. The calculus $CH$ is a process calculus where threads evaluate expressions from a lambda calculus extended by data constructors, case-expressions, recursive let-expressions, and Haskell’s seq-operator. Also monadic operations (sequencing and creating threads) are available. The shared memory is modeled by MVars (mutable variables) which are one-place buffers that can be either filled or empty and are blocking if a thread tries to fill a full buffer or to empty an empty buffer. The calculus $CH$ is also a variant (or a subcalculus) of the calculus $CHF$ [22,23] which extends Concurrent Haskell by...
futures in order to increase the declarativeness of such languages. Technically, we will reuse studies and results on the contextual semantics of CHF also for CH.

The contextual semantics of concurrent programming languages is a generalization of extensionality: The test for a program \( P \) is whether \( C[P] \) for all program-contexts \( C \) successfully terminates (converges) or not. For concurrent programs or processes we use two observations: may-convergence – at least one execution path terminates successfully, and should-convergence – every intermediate state may-converges. An alternative nondeterministic observation is must-convergence (all execution paths terminate successfully). The advantage of equivalence notions based on may- and should-convergence are their invariance under fairness restrictions, the preservation of deadlock-freedom, and the equivalence of busy-wait and wait-until behavior (see e.g. [29]). Showing equal expressivity of two (concurrent) calculi by a translation \( \tau \) is founded on correctness of \( \tau \): \( \tau \) leaves may- and should-convergence invariant. The property which is the gold-standard is adequacy (see Definition 4.4), which holds if \( \tau(P_1) \leq \tau(P_2) \) implies \( P_1 \leq P_2 \), for the contextual preorders and for all processes \( P_1, P_2 \). Full-abstractness, i.e. \( \tau(P_1) \leq \tau(P_2) \) iff \( P_1 \leq P_2 \) only holds in the rare case that the two calculi are more or less the same. In this paper we will study whether the \( \pi \)-calculus can be embedded into CH, which requires to find at least one very good translation. Our translation \( \tau_0 \) (defined and analyzed in Sect. 4) is one of these.

Our Approach. The main issue in finding a correct translation from \( \pi \)-processes to CH-programs is to correctly encode the synchronous communication of the \( \pi \)-calculus. Indeed, a considerable part of this paper will deal with this, and we will provide different proposals. The problem is that the so-called MVars in CH (see also [16]) have an asynchronous behavior. To implement synchronous communication the weaker synchronization property of MVars has to be exploited, where we must be aware of the potential interference of the executions of other translated communications on the same channel. The task of finding such translations is reminiscent of the channel-encoding used in [16], but, however, there an asynchronous channel is implemented while we look for synchronous communication.

We first provide a translation \( \tau_0 \) which uses a private MVar, which ensures that no other sender or receiver can interfere with the interaction. This idea was already used in [9,3]. We succeeded in mathematically confirming that the translation \( \tau_0 \) has strong and nice semantic properties.

Since we are also interested in simple translations and efficient results (for instance, the approach of using private names generates a new MVar for each communication), we tried to find correct translations with a fixed and static number of MVars per channel in the \( \pi \)-calculus. Since this task is too complex and error-prone for hand-crafting, we automated the search by implementing a procedure to rule out incorrect translations\(^3\). Thereby we assumed per commu-

\(^3\) The tool and some output generated by the tool are available via http://goethe.link/refute-pi
communication channel a single MVar for exchanging the channel-name and perhaps several additional MVars of unit type to perform checks whether the message was send / or received (we will call them check-MVars). The summarized results of this automated search technique are that there is a correct translation that uses two check-MVars, where one is used as a mutex between all senders or receivers on one channel, and there are correct translations using three additional MVars where the filling and emptying operations for each MVar must not come from the same sender or receiver. The experiments lead to the conjecture that there is no translation using only one check-MVar. We also provide formal arguments for the correctness of some of the minimal translations with global names.

These results will also permit to draw the conclusion that the expressiveness of the π-calculus is completely available in Concurrent Haskell.

Results. Our novel results are the definition of a translation \( \tau_0 \) from the π-calculus into CH, adequacy of the open translation \( \tau \) (Theorem 4.10), and full abstraction of \( \tau \) on closed π-processes (Theorem 4.12). From a technical point of view, a novelty is the comparison of the π-calculus with a concurrent programming language using contextual semantics for may-convergence and should-convergence in both calculi, which is technically involved since the syntactic details of the standard reductions in both calculi have to be analyzed. The adaptation of the adequacy and full abstraction notions (Definition 4.4) for open processes is also novel [27,28]. We further define a quite general formalism for the representation of translations with global names and analyze different classes of such translations by an automated tool. In particular, we show correctness of two particular of these translations in Theorems 5.9 and 5.11.

Further Related Work. Encodings of synchronous communication by asynchronous communication using a private name mechanism are given in [9,3] for (variants of the) π-calculus. Our idea of the translation \( \tau_0 \) similarly uses a private MVar to encode the channel based communication, but clearly our setting is different, since our target language is Concurrent Haskell. In [15] the expressiveness of synchronous and asynchronous communication in the π-calculus was compared where non-encodability is a main result, however, with several restrictions on the encoding. For instance rule out encodings that do not encoding parallel composition into parallel composition etc. We do not use such restrictions, and our approach focuses on the equivalence w.r.t. may- and should-convergence and the semantic notion of adequacy and full-abstractness.

Work on translations of the π-calculus into programming languages, calculi and logical systems and investigating the properties is [2], where a translation into a graph-rewriting calculus is given and soundness and completeness w.r.t. the operational behavior is proved. The article [30] shows a translation and a proof that the π-calculus is exactly operationally represented. There are several works on session types which are related to the π-calculus, for instance, [14] studies encodings from a session calculus into PCF extended by concurrency and effects and also an embedding in the other direction, mapping PCF extended by effects into a session calculus. The obtained result is a (quite strong) operational correspondence between both calculi.
Outline. The structure of the paper is as follows. After introducing both calculi in Sects. 2 and 3, the translation using private names is defined, explained and proved to be correct in Sect. 4. In Sect. 5 we introduce translations with global names and present our framework for automated search of (invalid) translations. We conclude and discuss future work in Sect. 6. To improve readability and due to space constraints several proofs are given in a technical appendix.

2 The π-Calculus with Stop

We explain the synchronous π-calculus without sums and with replication in a variant extended with a constant Stop, that signals successful termination of the whole π-calculus program. The π-calculus without Stop and with so-called barbed convergences are equivalent w.r.t. contextual semantics (see Theorem D.4 in the appendix). Thus, adding the constant Stop is not essential, however, the treatment and the translation are easier to explain for the π-calculus with Stop.

Definition 2.1. Let \( N \) be a countable set of (channel) names and \( x, y \in N \). The syntax of processes is \( P, Q \in \Pi_{\text{Stop}} := \nu x. P \mid \tau y. P \mid x(y). P \mid ! P \mid P \mid Q \mid 0 \mid \text{Stop} \). Free names \( FN(P) \), bound names \( BN(P) \), and \( \alpha \)-equivalence \( =_\alpha \) in \( \Pi_{\text{Stop}} \) are as usual in the π-calculus. A process \( P \) is closed if \( FN(P) = \emptyset \). Let \( \Pi_{\text{Stop}}^c \) denote the closed processes in \( \Pi_{\text{Stop}} \).

We briefly explain the language constructs. Name restriction \( \nu x. P \) restricts the scope of name \( x \) to process \( P \), \( P \mid 0 \) is the parallel composition of \( P \) and \( Q \), the process \( \tau y. P \) waits on channel \( y \) to output \( y \) over channel \( x \) and becoming \( P \) thereafter, the process \( x(y). P \) waits on channel \( x \) to receive input, and after receiving the input \( z \), the process turns into \( P[z/y] \) (where \( P[z/y] \) is the substitution of all free occurrences of name \( y \) by name \( z \) in process \( P \) ), the process \( ! P \) denotes the replication of process \( P \), i.e. it behaves like an infinite parallel combination of process \( P \) with itself, the process \( 0 \) is the silent process that does nothing, and \( \text{Stop} \) is a process constant that signals successful termination. We sometimes write \( x(y) \) instead of \( x(y) \) as well as \( \tau y.0 \).

Definition 2.2. The only reduction rule of the \( \Pi_{\text{Stop}} \)-calculus is the so-called interaction \( \overset{\text{ia}}{\rightarrow} \) which is defined as \( x(y). P \mid \tau z. Q \overset{\text{ia}}{\rightarrow} P[z/y] \mid 1 Q \).

Definition 2.3. Let \( P, Q, R \) be processes and \( x, y \) channel names. Structural congruence \( = \) is the least congruence satisfying:

\[
\begin{align*}
P & \equiv Q, \text{ if } P =_\alpha Q & P \mid 0 & \equiv P & \nu x. \text{Stop} & \equiv \text{Stop} \\
P \mid (Q \mid R) & \equiv (P \mid Q) \mid R & \nu x. 0 & \equiv 0 & \nu x, y. P & \equiv \nu y, x. P \\
\nu x. (P \mid Q) & \equiv P \mid \nu x. Q, \text{ if } x \notin FN(P) & P \mid Q & \equiv Q \mid P & 1P & \equiv P \mid 1P
\end{align*}
\]

Remark 2.4. We did not include “new” laws for structural congruences on the constant \( \text{Stop} \), like \( \text{Stop} \mid \text{Stop} \) equals \( \text{Stop} \), or even \( \text{Stop} \mid P \) equals \( \text{Stop} \). We did not want to include them, since this would require to develop a lot of theory.
known from the π-calculus without Stop again for the calculus with Stop. Our view is that Stop is a mechanism to have an easy notion for successful termination, that can be easily replaced by other similar notions (e.g. observing an open input or output as in barbed testing). However, it is easy to prove that those mentioned equations hold on the semantic level, e.g. the contextual equivalence Stop ⊢ P ≈ Stop holds (where ≈ is defined below in Definition 2.10).

As example, consider the communication that sends name y over channel x and then sends u over channel y:

\[(x(z).\overline{y}.u.0 | y(x).0) \overset{ia}{\longrightarrow} (\overline{y}[y/z] | y(x).0) \equiv (\overline{y}.u.0 | y(x).0) \overset{ia}{\longrightarrow} (0 | 0) \equiv 0\]

**Definition 2.5.** A process context \(C \in \mathcal{C}\) is a process that has a hole \([·]\) at one process position. They are defined by the grammar

\[C \in \mathcal{C} ::= [·] | \overline{x}(y).C | x(y).C | C|P | P|C | !C, \text{ with } x, y \in \mathbb{N}\]

With \(C[P]\) we denote the substitution of the hole in \(C\) by \(P\). Reduction contexts are defined as \(D \in \mathcal{PCtxt} \pi ::= [·] | D | P | P | D | !x.C, \text{ with } x, y \in \mathbb{N}\).

**Definition 2.6.** A standard reduction \(\overset{sr}{\longrightarrow}\) is the application of \(\overset{ia}{\longrightarrow}\) within a reduction context (modulo structural congruence):

\[P \equiv D[P'], P' \overset{ia}{\longrightarrow} Q', D[Q'] \equiv Q, \text{ and } D \in \mathcal{PCtxt}\]

\[P \overset{sr}{\longrightarrow} Q\]

Let \(\overset{sr,n}{\longrightarrow}\) denote \(n\) standard reductions and \(\overset{sr,\ast}{\longrightarrow}\) denotes the reflexive-transitive closure of \(\overset{sr}{\longrightarrow}\).

**Definition 2.7.** A process \(P \in \Pi_{\mathcal{Stop}}\) is successful, if \(P \equiv D[\text{Stop}]\) for some \(D \in \mathcal{PCtxt}\).

A property of interest is whether standard reductions successfully terminate or not. Since reduction is nondeterministic, we define observations which test the existence of a successful sequence (may-convergence), and which test all reduction possibilities (should-convergence):

**Definition 2.8.** Let \(P\) be a \(\Pi_{\mathcal{Stop}}\)-process. We say \(P\) is may-convergent (written \(P\downarrow\)), iff there is a successful process \(P'\) with \(P \overset{sr,\ast}{\longrightarrow} P'\). We say \(P\) is should-convergent (written \(P\downarrow\)), iff, for all \(P'\): \(P \overset{sr,\ast}{\longrightarrow} P'\) implies \(P'\downarrow\). If \(P\) is not may-convergent, then \(P\) is must-divergent (written \(P\uparrow\)). If \(P\) is not should-convergent, then we say it is may-divergent (written \(P\uparrow\)).

**Example 2.9.** The process \(P := \nu x, y. (x(z).0 | y(x).0)\) is may-convergent (\(P\downarrow\)) and should-convergent (\(P\downarrow\)), since \(P \overset{sr}{\longrightarrow} 0 | \text{Stop}\) is the only standard reduction sequence for \(P\). The process \(P' := \nu x, y. (x(z).0 | y(x).0)\) deterministically reduces to the silent process (i.e. \(P' \overset{sr}{\longrightarrow} 0\)), hence it is may-divergent (\(P'\uparrow\)) and
even must-divergent ($P' \uparrow$). The process $P'' := \nu x, y.(\exists y.0 \mid x(z).\text{Stop} \mid x(z).0)$ shows that may-convergence and should-convergence are different. There are two standard-reduction possibilities for $P''$: we have $P'' \xrightarrow{sr} \nu x, y.(\text{Stop} \mid x(z).0)$ and $P'' \xrightarrow{sr} \nu x, y.x(z).\text{Stop}$, where the first result is successful, and the second result is not successful. Hence, $P''$ is may-convergent but not should-convergent. It is also may-divergent; but not must-divergent.

Note that should-convergence implies may-convergence, and that must-divergence implies may-divergence.

**Definition 2.10.** For $P, Q \in \Pi_{\text{Stop}}$ and observation $\xi \in \{\downarrow, \downarrow, \uparrow, \uparrow\}$, we define $P \leq_c \xi Q$ iff $P \xrightarrow{\xi} Q$. The $\xi$-contextual preorders and $\xi$-contextual equivalences are defined as

$$P \leq_{c,\xi} Q \iff \forall C : C[P] \leq_{\xi} C[Q] \text{ and } P \sim_{c,\xi} Q \iff P \leq_{c,\xi} Q \land Q \leq_{c,\xi} P$$

**Lemma 2.11.** Let $P$ be a $\Pi_{\text{Stop}}$-process s.t. $FN(P) \subseteq \{x_1, \ldots, x_n\}$.

1. If $P \xrightarrow{sr} P'$ then $\nu x_1, \ldots, \nu x_n.P \xrightarrow{sr} \nu x_1, \ldots, \nu x_n.P'$.
2. If $\nu x_1, \ldots, \nu x_n.P \xrightarrow{sr} P'$ then $\nu x_1, \ldots, \nu x_n.P \equiv P' \equiv \nu x_1, \ldots, \nu x_n.P''$ and $P \xrightarrow{sr} P''$

**Lemma 2.12.** Let $\xi \in \{\downarrow, \uparrow, \downarrow, \uparrow\}$, $P, Q$ be $\Pi_{\text{Stop}}$-processes. Then $P \leq_{c,\xi} Q$ if, and only if $\forall C \in C$ such that $C[P]$ and $C[Q]$ are closed: $C[P] \leq_{\xi} C[Q]$.

**Proof.** One direction is trivial. For the other direction, we first consider $\xi = \downarrow$. Assume that $\forall C \in C$ such that $C[P]$ and $C[Q]$ are closed: $C[P] \leq_{\downarrow} C[Q]$. Let $C$ be an arbitrary context such that $C[P] \downarrow$ and let $FN(C[P] \mid C[Q]) = \{x_1, \ldots, x_n\}$. Lemma 2.11 shows $\nu x_1, \ldots, \nu x_n.C[P] \downarrow$. The precondition shows that $\nu x_1, \ldots, \nu x_n.C[Q] \downarrow$. Lemma 2.11 shows $C[Q] \downarrow$.

The case $\xi = \uparrow$ holds, since $\uparrow = \neg \downarrow$: If $\forall C \in C$ such that $C[P]$ and $C[Q]$ are closed: $C[P] \leq_{\uparrow} C[Q]$; then this shows $C[Q] \leq_{c,\uparrow} C[P]$ which is equivalent to $C[P] \leq_{c,\uparrow} C[Q]$.

For $\xi = \uparrow$, assume that $\forall C \in C$ such that $C[P]$ and $C[Q]$ are closed: $C[P] \leq_{\uparrow} C[Q]$. Let $C$ be an arbitrary context such that $C[P] \uparrow$ and let $FN(C[P] \mid C[Q]) = \{x_1, \ldots, x_n\}$. Lemma 2.11 and since adding or removing $\nu$-binders does not change success nor must-divergence we have $\nu x_1, \ldots, x_n.C[P] \uparrow$. Now the precondition shows $\nu x_1, \ldots, x_n.C[Q] \downarrow$, and Lemma 2.11 and since adding or removing $\nu$-binders does not change success nor must-divergence shows $C[Q] \uparrow$. Finally, the case $\xi = \uparrow$ follows by symmetry and since $P \uparrow \iff \neg(P \downarrow)$.

Since applying structural congruence is highly nondeterministic, and in order to facilitate reasoning on standard reduction sequences, we define a more restrictive use of the congruence laws which is still nondeterministic, but only applies the laws on the surface of the process.
Definition 2.13. Let $\text{dsc}_{\pi}$ be the union of the following rules, where $D \in \text{PCtxt}_{\pi}$:

$(\text{assocl}) \quad D[P | (Q | R)] \rightarrow D[(P | Q) | R] \quad (\text{assocr}) \quad D[(P | Q) | R] \rightarrow D[P | (Q | R)]$

$(\text{commute}) \quad D[P | Q] \rightarrow D[Q | P] \quad (\text{replunfold}) \quad D[!]P \rightarrow D[P | !P]$

$(\text{nuup1}) \quad D[(\nu z.P) | Q] \rightarrow D[\nu z.(P | Q)], \quad (\text{nuup2}) \quad D[\nu x.\nu z.P] \rightarrow D[\nu z.\nu x.P]$

if $z$ does not occur free in $Q$

Let $\text{dia} \rightarrow$ be the closure of $\text{isa}$ by reduction contexts $\text{PCtxt}_{\pi}$ and let $\text{dsr} \rightarrow$ be defined as the composition $\text{dia} \rightarrow \circ \text{dsr} \rightarrow \circ \text{dia} \rightarrow$.

We omit the proof of the following equivalences, but it can be constructed analogous to the proof given in [20] for barbed may- and should-testing. The theorem allows us to restrict standard reduction to $\text{dsr} \rightarrow$-reduction when reasoning on reduction sequences that witness may-convergence or may-divergence, resp.

Theorem 2.14. For all processes $P \in \Pi_{\text{Step}}$ the following holds:
1. $P \downarrow$ iff $P \text{dsr} \rightarrow \circ \text{Stop}$.
2. $P \uparrow$ iff $\exists P'$ such that $P \text{dsr} \rightarrow \circ P'$ and $P' \uparrow$.

3 The Process Calculus CH

We introduce the calculus $\text{CH}$ (which is a variant of the calculus $\text{CHF}$ in [22,23]) which models a core language of Concurrent Haskell [16]. We assume a partitioned set of data constructors $c$ where each family represents a type $T$. The data constructors of type $T$ are $c_{T,1}, \ldots, c_{T,|T|}$ where each $c_{T,i}$ has an arity $\text{ar}(c_{T,i}) \geq 0$. We assume that there is a type $\emptyset$ with data constructor $\emptyset$, a type $\text{Bool}$ with constructors $\text{True}, \text{False}$, a type $\text{List}$ with constructors $\text{Nil}$ and $\cdot$ (written infix as in Haskell), and a type $\text{Pair}$ with a constructor $(\cdot, \cdot)$ written as $(a, b)$. The syntax of the calculus $\text{CH}$ has processes $P \in \text{Proc}_{\text{CH}}$ on the top-layer which may have expressions $e \in \text{Expr}_{\text{CH}}$ as subterms. The syntax of both is shown in Fig. 1 where $u, w, x, y, z$ denote variables from a countably-infinite set of variables $\text{Var}$. As in the $\Pi_{\text{Step}}$-calculus, parallel processes are formed by parallel composition “$|$”. The $\nu$-binder restricts the scope of a variable. A concurrent thread $\nu e$ evaluates the expression $e$. In a process there is (at most one) unique distinguished thread, called the main thread written as $\text{main} e$. MVars are mutable variables which are empty or filled. A thread blocks if it wants to fill a filled MVar $x \text{m} e$ or empty an empty MVar $x \text{m} \cdot$. The variable $x$ is called the name of the MVar. Bindings $x = e$ model the global heap, where $x$ is called a binding variable. If variable $x$ is a name of an MVar or a binding variable, then $x$ is called an introduced variable. An introduced variable is visible to the whole process unless its scope is restricted by a $\nu$-binder, i.e. in $Q \nu x. P$ the scope of introduced variable $x$ is process $P$. A process is well-formed, if all introduced variables are pairwise distinct and there exists at most one main thread $\text{main} e$.

Expressions $\text{Expr}_{\text{CH}}$ consist of functional expressions and monadic expressions $\text{MExpr}_{\text{CH}}$ that model IO-operations. Functional expressions are variables,
\[ P \in \text{Proc}_{CH} ::= (P_1 \mid P_2) \mid e \mid \nu x. P \mid x \, m \, e \mid x \, m - \mid x = e \]
\[ e \in \text{Expr}_{CH} ::= x \mid \lambda x. e \mid (e_1 \mid e_2) \mid \text{seq} \, e_1 \, e_2 \mid e \, c_1 \ldots e_{c(e)} \mid \text{letrec} \, x_1 = e_1, \ldots, x_n = e_n \, \text{in} \, e \mid m \mid \text{case} \, T \, e \, \text{of} \, \big( e_T, x_1 \ldots x_{c(e_T)}, \rightarrow e_1 \big) \ldots \big( e_T, | \, x_1 \ldots x_{c(e_T)}, \rightarrow e_1 \big) \]
\[ m \in \text{MExpr}_{CH} ::= \text{return} \, e \mid e \gg \equiv e' \mid \text{fork} \, I \, O \, e \mid \text{takeMVar} \, e \mid \text{newMVar} \, e \mid \text{putMVar} \, e \, e' \]
\[ t \in \text{Typ}_{CH} ::= \mathsf{I} \, t \mid (T \, t_1 \ldots t_n) \mid \text{MVar} \, t \mid t_1 \rightarrow t_2 \]

Fig. 1. Syntax of expressions, processes, and types of \( CH \)

\[ P_1 \mid P_2 \equiv P_2 \mid P_1 \mid P_1 \mid P_2 \mid P_3 \]
\[ (P_1 \mid P_2) \mid P_3 \equiv P_3 \mid (P_2 \mid P_3) \]
\[ (\nu x. P) \mid P_2 \equiv \nu x. (P_1 \mid P_2), x \notin \text{FV}(P_2) \]
\[ \nu x_1. \nu x_2. P \equiv \nu x_2. \nu x_1. P \]
\[ P_1 \mid P_2 \text{ if } P_1 =_a P_2 \]

Fig. 2. Structural congruence of \( CH \)

Monadic Computations:

\begin{align*}
\text{(unit)} & \quad \equiv \mathbb{M}[\text{return} \, e_1] \gg \equiv e_2 \xrightarrow{\sigma_T} \equiv \mathbb{M}[e_2 \, e_1] \\
\text{(tvar)} & \quad \equiv \mathbb{M}[\text{takeMVar} \, x] \mid x \, m \, e \xrightarrow{\sigma_T} \equiv \mathbb{M}[\text{return} \, e] \mid x \, m - \\
\text{(pmvar)} & \quad \equiv \mathbb{M}[\text{putMVar} \, x \, e] \mid x \, m \xrightarrow{\sigma_T} \equiv \mathbb{M}[\text{return} \, ()] \mid x \, m \, e \\
\text{(nmvar)} & \quad \equiv \mathbb{M}[\text{newMVar} \, e] \xrightarrow{\sigma_T} \equiv \nu x. (\equiv \mathbb{M}[\text{return} \, x] \mid x \, m \, e) \\
\text{(fork)} & \quad \equiv \mathbb{M}[\text{fork} \, I \, O \, e] \xrightarrow{\sigma_T} \equiv \mathbb{M}[\text{return} \, ()] \mid I \equiv e
\end{align*}

Functional Evaluation:

\begin{align*}
\text{(cpce)} & \quad \equiv \mathbb{M}[F[x]] \mid x = e \xrightarrow{\sigma_T} \equiv \mathbb{M}[F[e]] \mid x = e \\
\text{(mkbinds)} & \quad \equiv \mathbb{M}[F[\text{letrec} \, x_1 = e_1, \ldots, x_n = e_n \, \text{in} \, e]] \\
& \quad \xrightarrow{\sigma_T} \equiv \nu x_1 \ldots x_n. (\mathbb{M}[F[e]] \mid x_1 = e_1 \ldots x_n = e_n) \\
\text{(beta)} & \quad \equiv \mathbb{M}[F[\lambda x. (e_1 \mid e_2)] \xrightarrow{\sigma_T} \equiv \mathbb{M}[F[e_1 \mid e_2/x]] \\
\text{(case)} & \quad \equiv \mathbb{M}[F[\text{case} \, T \, (c \, e_1 \ldots e_n) \, \text{of} \, \ldots (c \, y_1 \ldots y_n \rightarrow e_1) \ldots]] \\
& \quad \xrightarrow{\sigma_T} \equiv \mathbb{M}[F[e_1/y_1, \ldots, e_n/y_n]] \\
\text{(seq)} & \quad \equiv \mathbb{M}[F[(\text{seq} \, v \, e)] \xrightarrow{\sigma_T} \equiv \mathbb{M}[F[e]] \quad \text{where } v \text{ is a functional value}
\end{align*}

Closure:

If \( P_1 \equiv \mathbb{D}[P_1'], P_2 \equiv \mathbb{D}[P_2'], \) and \( P_1' \xrightarrow{\sigma_T} P_2' \) then \( P_1 \xrightarrow{\sigma_T} P_2. \)

Capture avoidance:

We assume capture avoiding reduction for all reduction rules.

Fig. 4. Standard reduction rules of \( CH \) (call-by-name-version)
abstractions \( \lambda x.e \), applications \((e_1 \ e_2)\), seq-expressions \((\text{seq} \ e_1 \ e_2)\), constructor applications \((c \ e_1 \ldots \ e_{\alpha(c)}\)\), letrec-expressions \((\text{letrec} \ x_1 = e_1,\ldots, x_n = e_n \ in \ e)\), and case-\(T\)-expressions for every type \(T\). We abbreviate case-expressions as case\(T\) \(e\) of alts where alts are the case-\(T\)-alternatives. The case-\(T\)-alternatives have exactly one alternative \(\langle c_{T,i} x_1 \ldots x_{\alpha(c_{T,i})} \rightarrow e_i \rangle\) for every constructor \(c_{T,i}\) of type \(T\), where \(x_1,\ldots,x_{\alpha(c_{T,i})}\) (occurring in the pattern \(c_{T,i} x_1 \ldots x_{\alpha(c_{T,i})}\)) are pairwise distinct variables that become bound with scope \(e_i\). We omit the type index \(T\) in case\(T\) if it is clear from the context. In \((\text{letrec} \ x_1 = e_1,\ldots, x_n = e_n \ in \ e)\) the variables \(x_1,\ldots,x_n\) are pairwise distinct and the bindings \(x_i = e_i\) are recursive, i.e. the scope of \(x_i\) is \(e_1,\ldots,e_n\) and \(e\). Monadic operators newMVar, takeMVar, and putMVar are used to create, to empty and to fill MVars, the “bind”-operator \(\gg\gg\) implements the sequential composition of IO-operations, the forkIO-operator performs thread creation, and return lifts expressions to monadic expressions.

Functional values are abstractions and constructor applications. If a monadic expression is of the form \((\text{newMVar} \ e)\), \((\text{takeMVar} \ e)\), \((\text{return} \ e)\), \((\text{forkIO} \ e)\), or \((\text{putMVar} \ e_1 \ e_2)\), then it is a monadic value. A value is either a functional or a monadic value.

Abstractions, letrec-expressions, case-\(T\)-alternatives, and \(\nu x. P\) introduce variable binders and thus this induces free and bound variables, \(\alpha\)-renaming, and \(\alpha\)-equivalence \(=_{\alpha}\). Let \(FV(P)\) (\(FV(e)\), resp.) be the free variables of process \(P\) (expression \(e\), resp.). We assume the distinct variable convention to hold: free variables are distinct from bound variables, and bound variables are pairwise distinct. We assume that reductions perform \(\alpha\)-renaming to obey this convention. Structural congruence \(\equiv\) of CH-processes is the least congruence satisfying the laws in Fig. 2. It allows to treat parallel composition as an associative-commutative operator, to shift \(\nu\)-binders to the top-level of processes and to \(\alpha\)-rename processes.

We assume that expressions and processes are well-typed according to a standard monomorphic type system: Since the typing rules are standard, we omit them, but explain the syntax of types (see Fig. 1): \((\text{IO} \ t)\) stands for a monadic action with return type \(t\), \((\text{MVar} \ t)\) stands for an MVar with content type \(t\), and \(t_1 \rightarrow t_2\) is a function type. We treat constructors like overloaded constants to use them in a polymorphic way.

We introduce a call-by-name small-step reduction for CH. This operational semantics can be shown to be equivalent to (a more realistic) call-by-need semantics. The proof is a copy of the proof given in [22] for the calculus CHF. However, the equivalence of the reduction strategies is not important or needed for the current paper. That is why we do not include it.

A context is a process or an expression with a (typed) hole \([\cdot]\). We introduce some classes of contexts in Fig. 3. On the process level there are process contexts \(PC\text{txt}_{\text{CH}}\), on expressions first monadic contexts \(MC\text{txt}_{\text{CH}}\) are used to find the next to-be-evaluated monadic action in a sequence of actions. For the evaluation of (purely functional) expressions, usual (call-by-name) expression evaluation contexts \(E \in EC\text{txt}_{\text{CH}}\) are used, and to enforce the evaluation of the (first)
argument of the monadic operators takeMVar and putMVar the class of forcing contexts \( \mathbb{F} \in FCxt_{CH} \) is used.

**Definition 3.1.** The standard reduction \( \rightarrow_s \) is defined by the rules and the closure in Fig. 4. We permit standard reduction only for well-formed processes which are not successful.

Functional evaluation includes \( \beta \)-reduction (beta), rule (cpe) for copying shared bindings into needed positions, rules (case) and (seq) to evaluate case- and seq-expressions, and rule (nmkbinds) to move letrec-bindings into the global set of shared bindings. For monadic computations, rule (lunit) implements the monadic evaluation by applying the first monad law. Rules (nmvar), (tmvar), and (pmvar) handle the MVar creation and access. A takeMVar-operation can only be performed on a filled MVar, and a putMVar-operation needs an empty MVar. Rule (fork) spawns a new concurrent thread. A concurrent thread finishes its computation if it is of the form \( \leftarrow \text{return } e \) (where \( e \) is of type \( \mathcal{O} \)).

A CH-process \( P \) is **successful** if \( P \equiv \nu x_1, \ldots, \nu x_n. (\text{\texttt{main return } e | P'}) \) and \( P \) is well-formed. These are the desired results of standard reduction sequences. The successful processes capture the behavior that termination of the main-thread implies termination of the whole program.

**Definition 3.2.** Let \( P \) be a CH-process. Process \( P \) may-converges (written as \( P \downarrow \)), iff \( P \) is well-formed and \( \exists P' : P \rightarrow_s P' \land P' \) successful. If \( P \downarrow \) does not hold, then \( P \) must-diverges written as \( P \uparrow \). Process \( P \) should-converges (written as \( P \uparrow \)), iff \( P \) is well-formed and \( \forall P' : P \rightarrow_s P' \Rightarrow P' \downarrow \). If \( P \) is not should-convergent, then we say \( P \) may-diverges written as \( P \uparrow \).

Note that a process \( P \) is may-divergent iff there is a finite reduction sequence \( P \rightarrow_s P' \) such that \( P' \uparrow \). Definition 3.2 implies that non-well-formed processes are always must-divergent, since they are irreducible and never successful.

**Definition 3.3.** Contextual approximation \( \leq_c \) and contextual equivalence \( \sim_c \) on CH-processes are defined as \( \leq_c := \leq_c, \downarrow \cap \leq_c, \uparrow \) and \( \sim_c := \leq_c \cap \geq_c \) where

\[
\begin{align*}
- P_1 \leq_c, \downarrow P_2 & \text{ iff } \forall D \in PCxt_{CH} : D[P_1] \downarrow \Rightarrow D[P_2] \downarrow \\
- P_1 \leq_c, \uparrow P_2 & \text{ iff } \forall D \in PCxt_{CH} : D[P_1] \uparrow \Rightarrow D[P_2] \uparrow
\end{align*}
\]

For CH-expressions, let \( e_1 \leq_c e_2 \) iff for all process-contexts \( C \) with a hole at expression position: \( C[e_1] \leq_c C[e_2] \) and \( e_1 \sim_c e_2 \) iff \( e_1 \leq_c e_2 \land e_2 \leq_c e_1 \).

The following equivalence will help to prove properties of our translation.

**Lemma 3.4.** The relations in Definition 3.3 are unchanged, if we add a closedness restriction as follows. Let \( \xi \in \{ \downarrow, \uparrow \} \), then \( P_1 \leq_c, \xi P_2 \) iff \( \forall D \in PCxt_{CH} \) such that \( D[P_1], D[P_2] \) are closed: \( D[P_1] \xi \Rightarrow D[P_2] \xi \).

**Proof.** One direction is obvious. For the other direction, let \( \xi \in \{ \downarrow, \uparrow \} \) and assume that for all \( D \) such that \( D[P_1], D[P_2] \) are closed: \( D[P_1] \xi \Rightarrow D[P_2] \xi \). Assume that \( D[P_1] \xi \) and \( V(D[P_1]) \cup V(D[P_2]) = \{ x_1, \ldots, x_n \} \). Since reductions
\[(gc) \forall x_1, \ldots, x_n. (P | \text{Comp}(x_1) | \ldots | \text{Comp}(x_n)) \rightarrow P,\]
if \(1 \leq i \leq n : x_i \notin \text{FV}(P)\) and \(\text{Comp}(x_i)\) is \(x_i = e_i, x_i \mathbf{m} e_i\), or \(x_i \mathbf{m} -\)
\[(\text{dtmvar}) \nu x. \mathbb{D}[\nu x. \mathbb{D}[\mathbb{M}[\text{takeMVar } x] | x \mathbf{m} e] | x \mathbf{m} -],\]
if \(\forall \nu x. \mathbb{D}[\mathbb{M}[\text{takeMVar } x] | x \mathbf{m} e] \rightarrow \nu x. \mathbb{D}[\mathbb{M}[\text{return } e] | x \mathbf{m} -],\)
\[(\text{dpmvar}) \nu x. \mathbb{D}[\mathbb{M}[\text{putMVar } x e] | x \mathbf{m} -] \rightarrow \nu x. \mathbb{D}[\mathbb{M}[\text{return } ()] | x \mathbf{m} e],\]
if \(\forall \nu x. \mathbb{D}[\mathbb{M}[\text{putMVar } x e] | x \mathbf{m} e] \rightarrow \nu x. \mathbb{D}[\mathbb{M}[\text{return } ()] | x \mathbf{m} e],\)
the first execution of \(\mathbb{M}[\text{takeMVar } x] | x \mathbf{m} e\) is in the shown thread.
\[(\text{dpmvar}) \nu x. \mathbb{D}[\mathbb{M}[\text{putMVar } x e] | x \mathbf{m} -] \rightarrow \nu x. \mathbb{D}[\mathbb{M}[\text{return } ()] | x \mathbf{m} -],\]
if \(\forall \nu x. \mathbb{D}[\mathbb{M}[\text{takeMVar } x] | x \mathbf{m} e] \rightarrow \nu x. \mathbb{D}[\mathbb{M}[\text{return } ()] | x \mathbf{m} e],\)
the first execution of \(\mathbb{M}[\text{putMVar } x e] | x \mathbf{m} e\) is in the shown thread.

**Fig. 5.** Garbage Collection and Deterministic Take und Put

are applicable with or without \(\nu\)-binders on the top, we have \(\nu x_1, \ldots, x_n. \mathbb{D}[P_1]| \xi\)
and by the precondition \(\nu x_1, \ldots, x_n. \mathbb{D}[P_2]| \xi\), since \(\nu x_1, \ldots, x_n. \mathbb{D}[\cdot]\) is a PCtxt-context. From \(\nu x_1, \ldots, x_n. \mathbb{D}[P_2]| \xi\) also \(\mathbb{D}[P_2]| \xi\) follows, since reductions are applicable with or without \(\nu\)-binders on the top. This shows \(P_1 \leq c_\xi P_2\).

**Proposition 3.5.** Let \(P_1, P_2\) be well-formed and \(P_1 \equiv P_2\). Then \(P_1 \sim c P_2\).

A program transformation \(\eta \subseteq (\text{Proc}_{\text{CH}} \times \text{Proc}_{\text{CH}})\) is correct iff \(P_1 \eta \eta P_2 \implies P_1 \sim c P_2\). In **Fig. 5** we define garbage collection \((gc)\) and deterministic variants of \((\text{tnvar})\) and \((\text{pnvar})\).

The following proposition holds, since correctness of transformations can be transferred from the calculus \(\text{CHF}\) by an embedding \(\iota : \text{CH} \rightarrow \text{CHF}\), such that \(\iota(P) \sim c_{\text{CHF}} \iota(P')\) implies \(P \sim c_{\text{CH}} P'\). See Appendix C.3 for the full proof.

**Proposition 3.6.** The rules \((\text{lunit})\), \((\text{nmvar})\), \((\text{fork})\), \((\text{cpce})\), \((\text{mkbinds})\), \((\text{beta})\), \((\text{case})\), \((\text{seq})\), \((\text{gc})\), \((\text{dtmvar})\), and \((\text{dpmvar})\) are correct program transformations.

## 4 The Translation \(\tau_0\) with Private Names

We present a translation that uses private names (i.e. new MVars) which uses similar ideas as the translation of the pi-calculus into an asynchronous pi-calculus in \[3\], which was an early example of using private names within a translation. Later in Sect.\[5\] we discuss translations and investigations on those translations which do not use private names.

We define a translation \(\tau_0\) mapping \(\Pi_{\text{step}}\)-processes to \(\text{CH}\)-processes. It uses the data type \text{Channel}, defined in Haskell-syntax as

\[
\text{data Channel }= \text{Channel (MVar (Channel (MVar () )))}.
\]

The type \text{Channel} is a recursive data type, which can be initialized with a \(\bot\)-expression. In the following, we abbreviate the expression \((\text{case } e \ of \ (\text{Channel } m \rightarrow m))\) as \((\text{unchan } e)\) and we use \(a \gg b\) as abbreviation for \(a \gg (\lambda _{-} . b)\) and also use
Haskell’s do-notation as abbreviation, where
\[
\begin{align*}
\text{do } \{ x \leftarrow e_1; e_2 \} &= e_1 \gg= \lambda x. (\text{do } \{ e_2 \}) \\
\text{do } \{(x,y) \leftarrow e_1; e_2 \} &= e_1 \gg= \lambda z. \text{casePair } z \text{ of } (x,y) \rightarrow (\text{do } \{ e_2 \}) \\
\text{do } \{ e_1; e_2 \} &= e_1 \gg \text{ (do } \{ e_2 \}) \\
\text{do } \{ e \} &= e
\end{align*}
\]

As a further abbreviation, we write \( y \leftarrow \text{newEmptyMVar} \) inside a do-block to abbreviate the sequence \( y \leftarrow \text{newMVar} \bot; \text{takeMVar} y \), where \( \bot \) is a must-divergent expression, for instance \( \text{letrec } x = x \text{ in } x \).

The translation from the \( \Pi_{\text{Stop}} \)-calculus is done by using one MVar per channel which contains a pair consisting of the (translated) name of the channel and a further MVar used for the synchronization. This MVar is private, such that only the sender and the receiver know it. Privacy is established by the sender: it creates a new MVar for every send operation.

Definition 4.1. We define the translation \( \tau_0 \) and its inner translation \( \tau \) from the \( \Pi_{\text{Stop}} \)-calculus into the CH-calculus in Fig. 6 as follows. For contexts, the translations are the same where the context hole is treated like a constant and translated as \( \tau([\cdot]) = [\cdot] \).

The translation \( \tau_0 \) generates a main-thread and an MVar \( \text{stop} \). The main thread is then waiting for the MVar \( \text{stop} \) to be emptied. The inner translation \( \tau \) translates the constructs and constants of the \( \Pi_{\text{Stop}} \)-calculus into CH-expressions.
Remark 4.2. Except for the main-thread, the translation $\tau_0$ generates a valid Concurrent Haskell-program, i.e. if we write $\tau_0(P) = \mu e.~`main = e`$, we can execute the translation in the Haskell-interpreter.

4.1 Properties of the Translation $\tau_0$

In this section we define properties of translations and motivate and present the plan for proving the results for the translations $\tau_0$ and $\tau$, which will be in Sect. 4.2 (and details also in Appendix A).

For the following definition of translation $\tau$ being compositional, adequate, or fully abstract, we adopt the view of the translation $\tau$ to be a translation from $\Pi_{\text{Stop}}$ into the $\text{CH}$-language with a special initial evaluation context $C_{\tau_{\text{out}}}$.

**Definition 4.3.** The context $C_{\tau_{\text{out}}}$ for translation $\tau_0$ is defined as

$$C_{\tau_{\text{out}}} = \nu f, \text{stop.} \mu e. \begin{array}{l} \text{do} \{ \text{stop} \leftarrow \text{newMVar}(); \text{forkIO} [\cdot]; \text{putMVar} \text{stop} \} \end{array}$$

We define variants of may- and should-convergence of expressions $e$ within $C_{\tau_{\text{out}}}$ in $\text{CH}$:

$$e \downarrow_0 \text{iff } C_{\tau_{\text{out}}}[e] \downarrow \text{ and } e \downarrow_0 \text{ iff } C_{\tau_{\text{out}}}[e] \downarrow.$$  

We define the relations $\leq_{c,\tau_0}$ and $\sim_{c,\tau_0}$ on $\text{CH}$-expressions:

$$e_1 \sim_{c,\tau_0} e_2 \text{ iff } e_1 \leq_{c,\tau_0} e_2 \text{ and } e_1 \leq_{c,\tau_0} e_2.$$  

Note that $\leq_{c,\text{CH}}$ is a subset of $\leq_{c,\tau_0}$, and hence also $\sim_{c,\text{CH}}$ is a subset of $\sim_{c,\tau_0}$ and thus we often can use the more general relations for reasoning and in particular can reuse results from $\text{CH}$.

**Definition 4.4.** Let $\Pi_{\text{Stop},C}$ be the contexts of $\Pi_{\text{Stop}}$. We define the following properties for $\tau_0$ and $\tau$ (see [27,28]). For open processes $P, P'$, we say that translation $\tau$ is

- **compositional upto** $\{\downarrow_0, \downarrow_0\}$ if for all $P \in \Pi_{\text{Stop}}$, all $C \in \Pi_{\text{Stop},C}$, and all $\xi \in \{\downarrow_0, \downarrow_0\}$ if $\text{FV}(C[P]) \subseteq \{\text{stop}\}$, then $\tau(C[P])\xi \Leftrightarrow \tau(C)\tau(P)\xi$,
- **adequate** if for all processes $P, P' \in \Pi_{\text{Stop}}$: $\tau(P) \leq_{c,\tau_0} \tau(P') \Rightarrow P \leq_c P'$, and
- **fully abstract** if for all processes $P, P' \in \Pi_{\text{Stop}}$: $P \leq_c P' \Leftrightarrow \tau(P) \leq_{c,\tau_0} \tau(P')$.

Convergence equivalence of the translation $\tau_0$ for may- and should-convergence holds. For readability and space reasons the proofs are omitted, but given in a technical appendix. In Appendix A we thus show the following two propositions:
Proposition 4.5. Let \( P \in \Pi_{\text{Stop}} \) be closed. Then \( \tau_0 \) is convergence-equivalent for \( \downarrow \), i.e. \( P \downarrow \) is equivalent to \( \tau_0(P) \downarrow \). This also implies that \( P \uparrow \) is equivalent to \( \tau(P) \uparrow \).

Proposition 4.6. Let \( P \in \Pi_{\text{Stop}} \) be closed. Then \( \tau_0 \) is convergence-equivalent for \( \downarrow \), i.e. \( P \downarrow \) is equivalent to \( \tau_0(P) \downarrow \).

We informally describe the main steps of the proof. Both propositions have an “if” and an “only-if” part, i.e. to prove them, one has to show four claims:

1. \( \tau_0 \) preserves may-convergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( P \downarrow \Rightarrow \tau_0(P) \downarrow \).
2. \( \tau_0 \) reflects may-convergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( \tau_0(P) \downarrow \Rightarrow P \downarrow \).
3. \( \tau_0 \) preserves must-convergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( P \uparrow \Rightarrow \tau_0(P) \uparrow \).
4. \( \tau_0 \) reflects must-convergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( \tau_0(P) \uparrow \Rightarrow P \uparrow \).

Actually, for the latter two parts, we show reflection and preservation of may-divergence, i.e.:

3’. \( \tau_0 \) reflects may-divergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( \tau_0(P) \uparrow \Rightarrow P \uparrow \).
4’. \( \tau_0 \) preserves may-divergence, i.e. for closed \( P \in \Pi_{\text{Stop}} \): \( P \uparrow \Rightarrow \tau_0(P) \uparrow \).

Note that claim 3 is equivalent to 3’ and claim 4 is equivalent to claim 4’.

The proof technique for showing parts 1,2,3’,4’ is to investigate properties of reduction sequences (those that end in a successful process and those that end in a must-divergent process) and to (inductively) apply appropriate (and permitted) rearrangements of reduction sequences and correct program transformations, i.e. those listed in Proposition 3.6.

For the given reduction sequences in the \( \Pi_{\text{Stop}} \)-calculus (giving evidence that \( P \downarrow \), or \( P \uparrow \) resp. holds) for proving part 1 and 4’, we assume that they are given as steps of \( dsr \)-reductions (Definition 2.13). This simplifies the proof and is correct due to Theorem 2.14.

A further technical detail is that we use another translation \( \sigma \) (and \( \sigma_0 \)) which translates several \( \pi \)-process components directly into CH-process components (for instance, parallel composition is directly translated into parallel composition), instead of translating them into code which generates the components (as \( \tau \) does). An intermediate step in the proof is to show equivalence of \( \sigma \) and \( \tau \) w.r.t. contextual equivalence.

For proving the reflection parts, i.e. part 2 and part 3, the following diagrams sketch the overall idea of the induction proof, which mainly is on the number \( n \) of given reductions.

For part 2 the diagram is

\[
\begin{array}{ccc}
P \xrightarrow{sr,\ast} P_1 & \xrightarrow{\sigma_0} & \sigma_0(P) \\
\end{array}
\]

For part 3, the diagram is
The difference of both parts is the induction base: For part 2, the processes are successful, while for part 3, the processes have to be must-divergent. Both diagrams also illustrate the use of correct program transformations to obtain a “reordered” sequence of reductions and transformation for \( \sigma_0(P) \) such that it can be back-translated into a sequence of \( sr \)-reductions in the \( II_{\text{Stop}} \)-calculus.

### 4.2 Main Results for the Translation \( \tau_0 \)

We show the main result that the translation is adequate (Theorem 4.10). The interpretation of this result is that the pi-calculus with the concurrent semantics is semantically represented within \( CH \). This result is on a more abstract level, since it is based on the property whether the programs (in all contexts) produce values or may run into failure, or get stuck; or not. Since the pi-calculus does not have a notion of values, like numbers or lists, also the translated processes cannot be compared w.r.t. values other than a single form of value.

The translation \( \tau_0 \) is not fully abstract (Theorem 4.11\(^*\)), which is rather natural, since it only means that it is mapped into a subset of the expressions and that this is a proper subset w.r.t. the semantics.

First we show a simple form of a context lemma:

**Lemma 4.7.** Let \( e, e' \) be \( CH \)-expressions, where the only free variable in \( e, e' \) is stop. Then \( C^\tau_{\text{out}}[e] \leq_c C^\tau_{\text{out}}[e'] \) iff \( C^\tau_{\text{out}}[e] \downarrow \implies C^\tau_{\text{out}}[e'] \downarrow \) and \( C^\tau_{\text{out}}[e] \uparrow \implies C^\tau_{\text{out}}[e'] \uparrow \).

**Proof.** One direction is obvious by using the empty context in the \( \leq_c \)-definition. For the other direction, assume \( C^\tau_{\text{out}}[e] \downarrow \implies C^\tau_{\text{out}}[e'] \downarrow \) and \( C^\tau_{\text{out}}[e] \uparrow \implies C^\tau_{\text{out}}[e'] \uparrow \). There are two cases:

1. Let \( D \) be a process-context such that \( D[C^\tau_{\text{out}}[e]] \) and \( D[C^\tau_{\text{out}}[e']] \) are closed, and \( D[C^\tau_{\text{out}}[e]] \downarrow \). Then we have to show that \( D[C^\tau_{\text{out}}[e']] \downarrow \). Due to closedness, \( D[C^\tau_{\text{out}}[e]] \equiv P_{D} \mid C^\tau_{\text{out}}[e] \) and \( D[C^\tau_{\text{out}}[e']] \equiv P_{D} \mid C^\tau_{\text{out}}[e'] \) for some closed \( CH \)-process \( P_{D} \). Due to closedness, there is no interference between \( P_{D} \) and \( C^\tau_{\text{out}}[e] \), or \( C^\tau_{\text{out}}[e'] \), resp. Hence \( C^\tau_{\text{out}}[e] \downarrow \implies C^\tau_{\text{out}}[e'] \downarrow \), and thus \( D[C^\tau_{\text{out}}[e']] \downarrow \).

2. Assume that \( D[C^\tau_{\text{out}}[e']] \uparrow \). We have to show that \( D[C^\tau_{\text{out}}[e']] \uparrow \) holds. Due to closedness, we have \( D[C^\tau_{\text{out}}[e]] \equiv P_{D} \mid C^\tau_{\text{out}}[e] \) and \( D[C^\tau_{\text{out}}[e']] \equiv P_{D} \mid C^\tau_{\text{out}}[Q'] \) and the same arguments as in item 1 show the claim.

**Proposition 4.8.** The translation \( \tau \) is compositional up to \( \{ \downarrow_0, \uparrow_0 \} \).

**Proof.** This follows by checking whether the single cases of the translation \( \tau \) are independent of the surrounding context, and translate every level independently.
Theorem 4.9. Let $P \in \Pi_{\text{Stop}}$ be closed. Then $\tau_0$ is convergence-equivalent for \( \downarrow \) and \( \parallel \), i.e. $P \downarrow$ is equivalent to $\tau_0(P) \downarrow$, and $P \parallel$ is equivalent to $\tau_0(P) \parallel$. This also shows convergence-equivalence of $\tau$ w.r.t. $\downarrow_0$, $\parallel_0$, i.e. For closed $P \in \Pi_{\text{Stop}}$: $P \downarrow \iff \tau(P) \downarrow_0$ and $P \parallel \iff \tau(P) \parallel_0$.

Proof. This follows from Proposition 4.5 for may-convergence and from Proposition 4.6 for should-convergence.

We show in the following that the translation $\tau$ transports $\Pi_{\text{Stop}}$-processes into $\text{CH}$, such that adequacy holds. This is a rather strong statement: It shows that the translated processes also mimic the behavior of the original $\Pi_{\text{Stop}}$-processes when plugged into contexts in a correct way. If the translated open processes cannot be distinguished by $\leq_{\text{c},\tau_0}$, i.e. there is no test that detects a difference where may and should-convergence are the observables, then the original processes are equivalent in the pi-calculus.

However, this open translation is not fully abstract, which means that there are $\text{CH}$-contexts that can see and exploit too much of the details of the translation.

We state and prove the main result of our paper.

Main Theorem 4.10 The translation $\tau$ is adequate.

Proof. We prove the adequacy-property for the preorder $\leq_{\text{c},\tau_0}$. The property for the contextual equivalence $\sim_{\text{c},\tau_0}$ and $\sim$, then follows by symmetry. Let $P, P'$ be $\Pi_{\text{Stop}}$-processes, such that $\tau(P) \leq_{\text{c},\tau_0} \tau(P')$. We show that $P \leq_{\text{c}} P'$. Now we use the equivalence in Lemma 3.4 and restrict considerations to closed $C[P], C[P']$ below: Let $C$ be a context in $\Pi_{\text{Stop}}$, such that $C[P], C[P']$ are closed and $C[P] \parallel$. Then $\tau_0(C[P]) = C'_{\text{out}}[\tau(C[P])]$. Closed convergence equivalence shows that $C'_{\text{out}}[\tau(C[P])] \downarrow$, which is the same as $C'_{\text{out}}[\tau(C)[\tau(P)]] \downarrow$ by Proposition 4.8. The assumption $\tau(P) \leq_{\text{c},\tau_0} \tau(P')$ implies $C'_{\text{out}}[\tau(C)[\tau(P)]] \downarrow$, which again is the same as $C'_{\text{out}}[\tau(C'[P'])] \downarrow$ using Proposition 4.8. Again, closed convergence equivalence implies $C'[P'] \downarrow$. The same arguments hold for $\parallel$ instead of $\downarrow$. The computation can be done for every context $C$ that satisfies the conditions.

In summary, we obtain $P \leq_{\text{c}} P'$.

Theorem 4.11. The translation $\tau$ is not fully abstract.

Proof. This holds, since an open translation can be closed in $\text{CH}$ by a context without initializing the $\nu$-bound MVars. For $P = \bar{x}(y)\text{Stop} | x(z)\text{Stop}$, we have $P \sim \text{Stop}$ in the $\pi$-calculus (see Example D.2), but $\tau(P) \not\sim_{\text{c},\tau_0} \tau(\text{Stop})$ in $\text{CH}$: let $\bar{D}$ be a process context that does not initialize the MVars for $x$ (as the translation does). Then $\bar{D}[\tau(P)] \parallel_0$, but $\bar{D}[\tau(\text{Stop})] \parallel_0$.

However, restricted to closed processes, full abstraction holds:

Theorem 4.12. For closed $P_1, P_2 \in \Pi_{\text{Stop}}$, the equivalence $P_1 \leq_{\text{c}} P_2 \iff \tau(P_1) \leq_{\text{c},\tau_0} \tau(P_2)$ holds.

Proof. For closed $P, P'$, the implication $P \leq_{\text{c}} P' \implies \tau(P_1) \leq_{\text{c},\tau_0} \tau(P_2)$ follows from Lemma 4.7 since $\tau_0$ produces closed processes that are in context $C'_{\text{out}}$. The other direction follows from Main Theorem 4.10.
5 Translations with Global Names

In this section, we discuss translations that do not use private names, but only global names. We report on an automated tool that we implemented and that searches for those translations trying to refute their correctness. We only consider the aspect of how to encode the synchronous message passing of the π-calculus, the other aspects (encoding parallel composition, replication and the Stop-constant) are not discussed, where we assume that they are encoded as before. We also keep the main idea to translate a channel of the π-calculus into CH: we represent a π-channel in CH as an object of a user-defined data type Channel that consists of

1. an MVar for transferring the message (which again is a Channel),
2. additional MVars for implementing a correct synchronization mechanism.

In the translation τ in the previous section, we used a private MVar (which was created by the sender for every communication action and transferred together with the message). In this section we investigate translations where this mechanism is replaced by one or several public MVars, which are created once together with the channel object. To restrict the search space for translations, only the synchronization mechanism of MVars (by emptying and filling them) is used, but we forbid to transfer specific data (like numbers etc.). Hence, we restrict these MVars (which are called check-MVars from now on) to be of type MVar ()

In summary, we now analyze translations of π-calculus channels into a CH-data type Channel where the data type definition in Haskell is of the form

\[
\text{data Channel} = \text{Channel} \ (\text{MVar Channel}) (\text{MVar ()}) \ldots (\text{MVar ()}) \quad n \text{ times}
\]

A channel \( x \) of the π-calculus is then represented as a CH-binding

\[
x = \text{Channel content check}_1 \ldots \text{check}_n
\]

where content, check\(_1\), \ldots, check\(_n\) are appropriately initialized (i.e. empty) MVars. The MVars are not private (but public (or global), since all processes which know the name \( x \) have access to the \( n + 1 \) components of the channel.

After fixing this representation of a π-channel in CH, the task is to translate the input- and output actions \( x(y), \pi z \) as CH-programs such that the interaction reduction is performed correct and synchronously\(^4\).

Let us call the translation of \( x(y) \), the receiver program (the receiver, for short), and the translation of \( \pi z \) the sender program (the sender, for short). To simplify the analysis, we restrict the allowed operations of the sender and the receiver, and thus allow only the following operations:

\(^4\) We omit the translation of the \( \nu \)-operator at this moment, but clearly, it has to construct and initialize the appropriate MVars. In Definition 5.2, the full translation will be given.
put$S$: The sender puts its message into the contents-MVar of the channel. This operation occurs exactly once in the sender program. We write it as $\text{put}_S x z$, or even more abstractly as $\text{put}_S$, if $x$ and $z$ are clear from the context. It represents the CH-expression \( \text{case } x \text{ of } (\text{Channel } c a_1 \ldots a_n \rightarrow \text{putMVar } c z \gg e) \) where $e$ is the remaining program of the sender.

$\text{take}_S$: The receiver takes the message from the contents-MVar of channel $x$ and replaces the channel name $y$ by the received name in the subsequent program (which does not belong the current input-operation). This operation occurs exactly once in the receive program. We write this as $\text{take}_S x y$, or even more abstractly as $\text{take}_S$, if $x$ and $y$ are clear from the context. It represents the CH-expression

\[
\text{case } x \text{ of } (\text{Channel } c a_1 \ldots a_n \rightarrow \text{takeMVar } c \gg \lambda y.e)
\]

where $e$ is the remaining program of the receiver.

In do-notation, we also write $\text{do} \{ y \leftarrow \text{take}_S x; e \}$ to abbreviate the above CH-expression.

$\text{put}_C$ and $\text{take}_C$: The sender and the receiver may synchronize on a check-MVar $\text{check}_1, \ldots, \text{check}_n$ by either putting () into it or by emptying the MVar. These operations are written as $\text{put}_C^i$ and $\text{take}_C^i$, or even more abstractly as $\text{put}_C^i, \text{take}_C^i$ if the name $x$ is clear from the context. We write $\text{put}_C$ and $\text{take}_C$ if there is only one check-MVar.

Let $e$ be the remaining program of the sender or receiver. Then $\text{put}_C^i_x$ represents the CH-program

\[
\text{case } x \text{ of } (\text{Channel } c a_1 \ldots a_n \rightarrow \text{putMVar } a_i () \gg e)
\]

and $\text{take}_C^i$ represents the CH-program

\[
\text{case } x \text{ of } (\text{Channel } c a_1 \ldots a_n \rightarrow \text{takeMVar } a_i \gg e)
\]

We restrict our search for translations to the case that the sender and the receiver programs are sequences of the above operations, and that they are independent of the channel name $x$. With this restriction we can abstractly write the translation of the sender and the receiver program as a pair of sequences, where only $\text{put}_S, \text{take}_S, \text{put}_C^i$ and $\text{take}_C^i$ operations are used. We make some more restrictions in the following definition:

**Definition 5.1.** Let $n > 0$ be a number of check-MVars. A standard global synchronized-to-buffer translation (or gstb-translation) is represented as a pair $(T_{send}, T_{receive})$ of a send-sequence $T_{send}$ and of a receive-sequence $T_{receive}$ which both consist of $\text{put}_S, \text{take}_S, \text{put}_C^i$ and $\text{take}_C^i$ operations, where

- The send-sequence contains $\text{put}_S$ (exactly once), and the receive-sequence contains $\text{take}_S$ (exactly once),
- For every $\text{put}_C^i$-action in $(T_{send}, T_{receive})$, there is also a $\text{take}_C^i$-action in $(T_{send}, T_{receive})$. 
We can assume that in the send-sequence the indices \( i \) are ascending. I.e. if \( \text{putC}_i \) or \( \text{takeC}_i \) is before \( \text{putC}_j \) or \( \text{takeC}_j \), then \( i < j \).

In the following we often say translation instead of gstb-translation, if this is clear from the context.

**Definition 5.2.** Let \( (T_{send}, T_{receive}) \) be a gstb-translation. We write \( T_{x,y} \) for the program \( T_{send} \) instantiated for the output-prefix \( xy \), i.e. the \( \text{putS} \)-operation is \( \text{putS}_x y \), and all other operations are indexed with name \( x \), i.e \( \text{takeC}_i \) becomes \( \text{takeC}_i x \) and \( \text{putC}_i \) becomes \( \text{putC}_i x \). We write \( T_{x,y} \) for the program \( T_{receive} \) instantiated for the input-prefix \( xy \), i.e. the \( \text{takeS} \)-operation is \( \text{takeS}_x y \), and all other operations are indexed with name \( x \), i.e \( \text{takeC}_i \) becomes \( \text{takeC}_i x \) and \( \text{putC}_i \) becomes \( \text{putC}_i x \).

The induced translations \( \phi_0, \tau \) and \( \phi, \tau \) of \( (T_{send}, T_{receive}) \) are defined as follows, where \( n \) is the number of check-MVars used by the translation:

\[
\begin{align*}
\phi_0(P) &= \text{main} \leftarrow \text{do}\{ \text{stop} \leftarrow \text{newMVar}(); \text{forkIO } \phi(P); \text{putMVar stop} () \} \\
\phi(P_{xy}) &= \text{do}\{ T_{send}; \phi(P) \} \\
\phi(T_{x,y})(P) &= \text{do}\{ T_{receive}; \phi(P) \} \\
\phi(T_{x,y})(P | Q) &= \text{do}\{ \text{forkIO } \phi(Q); \phi(P) \} \\
\phi(T_{x,y})(\nu x.P) &= \text{do}\{ \text{contx} \leftarrow \text{newEmptyMVar}; \text{checkx}_1 \leftarrow \text{newEmptyMVar}; \ldots; \text{checkx}_n \leftarrow \text{newEmptyMVar}; \text{letrec } x = \text{Channel contx checkx}_1 \ldots \text{checkx}_n \text{ in } \phi(P) \} \\
\phi(P(0)) &= \text{return} () \\
\phi(P(\text{Stop})) &= \text{takeMVar stop} \\
\phi(P(!P)) &= \text{letrec } f = \text{do}\{ \text{forkIO } \phi(P); f \} \text{ in } f
\end{align*}
\]

Note that the induced translations are defined similar to the translation \( \tau_0 \) and \( \tau \), where the differences are the representations of the channel, and thus the translation of \( \nu x, y(P) \), and \( \pi y \) is changed, while the other cases remain the same. Especially, it holds that \( \phi_0(P) = C^{\tau_0}_{\text{out}}[\phi(P)] \). Using the same arguments as in the proof of Main Theorem 4.10 we can show the following proposition.

**Proposition 5.3.** If \( \phi \) is closed convergence equivalent, then \( \phi \) is adequate and on closed processes it is fully-abstract.

We also speak of an execution of a translation \( (T_{send}, T_{receive}) \) for name \( x \). Here we mean the simulation of the abstract program, i.e. a program that starts with empty MVars \( x, x_1, \ldots, x_n \), and is an interleaved sequence of actions from the send and receive-sequence \( T_{send} \) and \( T_{receive} \), respectively.

### 5.1 Classifying Translations

To speak about the translations we make some more classifications:

**Number \( n \) of check-MVars** We classify the translations by the number of check-MVars that are present in the Channel-type.
Multiple-Uses We say that a translation allows multiple uses, if the same check-MVar is used more than once, i.e. the sender and/or receiver may contain takeC\textsubscript{i} and putC\textsubscript{i} more than once for the same i.

Interprocess check restriction We say that a translation has the interprocess check restriction, if for every i: takeC\textsubscript{i} and putC\textsubscript{i} do not occur both in T\textsubscript{send}, and also not in T\textsubscript{receive}.

We define some properties of gshb-translations:

**Definition 5.4.** A translation (T\textsubscript{send}, T\textsubscript{receive}) according to Definition 5.1 is

– executable, if there is a deadlock free interleaving execution of the parallel combination of T\textsubscript{send} and T\textsubscript{receive}.

– communicating, if T\textsubscript{send} contains at least one takeC\textsubscript{i}-action.

– overlap-free, if for a fixed name x, starting with empty MVars, every interleaved (concurrent) execution of T\textsubscript{send} and T\textsubscript{receive} cannot be disturbed by starting another T\textsubscript{send} and T\textsubscript{receive}, after T\textsubscript{send} and T\textsubscript{receive} both performed (at least) one action.

5.2 Simulating Translations to Refute their Correctness

We implemented a tool to enumerate translations and then to test whether every translation preserves and reflects may- and should-convergence. The above mentioned parallel execution of T\textsubscript{send} and T\textsubscript{receive} is not sufficient to refute most of the translations, since it corresponds to the evaluation of the \(\pi\)-program \(\nu x.(x(y)\mid z)\) (which is must-divergent, since no Stop occurs). Thus our approach is to apply the translation to a subset of \(\pi\)-process for which we can decide may- and should-convergence (before and after the translation) in an automated way. Thus we consider only \(\pi\)-programs of the form \((\nu x_1, \ldots, x_n.P)\) where P contains only 0, Stop, parallel composition |, and input- and output-prefixes. Thus, the programs are replication free and the \(\nu\)-binders are on the top, and hence terminate. In the following we omit the \(\nu\)-binders, but always mean them to be present.

Our simulation tool\footnote{Available via \url{http://goethe.link/refute-pi}} can execute all possible evaluations of such \(\pi\)-processes and – since all evaluation paths are of finite length, the tool can check for may- and should-convergence of the \(\pi\)-program. For the translated program, we do not generate a full CH-program: We generate a sequence of sequences of takeS\textsubscript{i}, putS\textsubscript{i}, takeC\textsubscript{i}, putC\textsubscript{i} z and Stop operations by applying the translation to all action prefixes in \(\pi\)-program and by translating Stop into Stop, 0 into an empty sequence. We get a sequence of sequences, since we have several threads and each thread is represented by such a sequence. For executing the translated program we simulate the global store (of MVars) and execute all possible interleavings where we check for may- and should-convergence by looking whether the Stop eventually occurs at the beginning of the sequence. This quite obviously simulates the behavior of the real CH-program but in a controllable manner.
Now having an encoding of the sender- and the receiver program, we use a
π-calculus process $P$ and
1. Translate it with the encodings in the sequence of sequences consisting of $\text{takeS}_x, \text{putS}_x, \text{takeC}_x, \text{putC}_x$ and $\text{Stop}$-operations.
2. Simulate the execution on all interleavings
3. Test may- and should convergence of the original π-program $P$ as well as
the encoded program (w.r.t. the simulation)
4. Compare the convergence before and after the translation. If there is a
difference in the convergence behavior, then $P$ is a counter-example for the
correctness of the encodings.

5.3 Translations Without Check-MVars
The (trivial) case that there is no check-MVar leads to one possible translation
$([\text{putS}], [\text{takeS}])$ which means that $\nu x.z$ is translated into $\text{putS}_x y$ and $x(y)$ is
translated into $\text{takeS}_x y$. This translation is not correct, since for instance the π-
process $\nu x.x.\text{Stop}(y(x)).\text{Stop}$ is neither may- nor should-convergent, but the translation
$\text{putS}_x z; \text{takeS}_x y; \text{Stop}$ is may- and should-convergent, since the put- and the
take-operation can be done sequentially.

5.4 Translations With Interprocess Check Restriction
We first consider translations with the interprocess check restriction. In this case
the number of different translations is $n! \times 2^n \times (n + 1)^2$ (where $n$ is the number
of check-MVars), which for $n = 1$ results in 8, for $n = 2$ results in 72, and for
$n = 3$ in 768 translations.

For a single check-MVar all 8 translations are rejected by our simulation,
Table 1 shows the translations and the obtained counter examples.

Remark 5.5. We exhibit an example, why the translation
$$(T_{\text{send}}, T_{\text{receive}}) = ([\text{putS}, \text{takeC}], [\text{takeS}, \text{putC}])$$
is not correct. Consider the following $\Pi_{\text{Stop}}$-process
$$P = \nu x, z, w, a. (\nu z. \text{Stop} \ | \ \nu w. \text{Stop} \ | \ x(y). y(w). 0)$$
which is should-convergent, since the two possible reduction sequences starting
with $P$ end in successful processes.

Consider the translated CH-program where we use the commands abbreviating the concrete CH-programs, and where we assume that the program code
for creating the MVars and the bindings for the Channel-objects are already executed:

$$\nu \ldots . x = \text{Channel } c_x \text{ chk}_x | z = \text{Channel } c_z \text{ chk}_z | w = \text{Channel } c_w \text{ chk}_w | \ldots$$
$$| c_x.m - | \text{chk}_x.m - | c_z.m - | \text{chk}_z.m - | c_w.m - | \text{chk}_w.m - | \ldots$$
$$| \text{do} \{ \text{putS}_z ; \text{takeC}_z ; \text{putS}_z ; \text{takeC}_z ; \ldots \}$$
$$| \text{do} \{ \text{putS}_w ; \text{takeC}_x ; \text{putS}_w ; \text{takeC}_w ; \ldots \}$$
$$| \text{do} \{ y \leftarrow \text{takeS}_x ; \text{putC}_x ; w \leftarrow \text{takeS}_y ; \ldots \}$$
This reduces after executing putS\_x z and y ← takeS\_x

\[\nu.x = \text{Channel } c_x \text{ } \text{chk}_x | z = \text{Channel } c_z \text{ } \text{chk}_z | w = \text{Channel } c_w \text{ } \text{chk}_w | \ldots\]

\[| c_y m \leadsto | \text{chk}_x m \leadsto | c_z m \leadsto | \text{chk}_z m \leadsto | c_w m \leadsto | \text{chk}_w m \leadsto | \ldots\]

| do{takeC\_z; putS\_z a; takeC\_z; \ldots} |
| do{putS\_x w; takeC\_x; putS\_w a; takeC\_w; \ldots} |
| do{putC\_z; w ← takeS\_z; \ldots} |

The second and third thread make steps:

\[\nu.x = \text{Channel } c_x \text{ } \text{chk}_x | z = \text{Channel } c_z \text{ } \text{chk}_z | w = \text{Channel } c_w \text{ } \text{chk}_w | \ldots\]

\[| c_y m \leadsto | \text{chk}_x m | c_z m \leadsto | \text{chk}_z m | c_w m \leadsto | \text{chk}_w m \leadsto | \ldots\]

| do{takeC\_z; putS\_z a; takeC\_z; \ldots} |
| do{takeC\_z; putS\_w a; takeC\_w; \ldots} |
| do{w ← takeS\_z; \ldots} |

The second thread can make an unexpected step:

\[\nu.x = \text{Channel } c_x \text{ } \text{chk}_x | z = \text{Channel } c_z \text{ } \text{chk}_z | w = \text{Channel } c_w \text{ } \text{chk}_w | \ldots\]

\[| c_y m \leadsto | \text{chk}_x m | c_z m \leadsto | \text{chk}_z m | c_w m \leadsto | \text{chk}_w m \leadsto | \ldots\]

| do{takeC\_z; putS\_a; takeC\_z; \ldots} |
| do{putS\_a; takeC\_w; \ldots} |
| do{w ← takeS\_z; \ldots} |

Now the only possible step is:

\[\nu.x = \text{Channel } c_x \text{ } \text{chk}_x | z = \text{Channel } c_z \text{ } \text{chk}_z | w = \text{Channel } c_w \text{ } \text{chk}_w | \ldots\]

\[| c_y m \leadsto | \text{chk}_x m | c_z m \leadsto | \text{chk}_z m | c_w m \leadsto | \text{chk}_w m \leadsto | \ldots\]

| do{takeC\_z; putS\_a; takeC\_z; \ldots} |
| do{takeC\_w; \ldots} |
| do{w ← takeS\_z; \ldots} |

and now the process is stuck.
If we use 2 check-MVars, then there are 72 translations. Again our tool is able to refute all of them. Compared to Table 1, there are two further \( \pi \)-processes used, which act as counterexample. However, one can also use the processes \( x(y). \text{Stop} \) | \( x(y) \) and \( x(z)y(u). \text{Stop} \) to refute all 72 translations.

For 3 MVars, there are 768 translations, where 762 are rejected (where the counter examples are the same as for 2-check-MVars) and 6 potentially correct translations remain. The six translations are shown in Table 2.

<table>
<thead>
<tr>
<th>Sender</th>
<th>Receiver</th>
</tr>
</thead>
<tbody>
<tr>
<td>([\text{putS, putC}, \text{takeC}], \text{putC} )</td>
<td>([\text{takeC}, \text{putC}, \text{takeC}, \text{takeS}] )</td>
</tr>
<tr>
<td>([\text{takeC}, \text{putS}, \text{takeC}, \text{takeC}] )</td>
<td>([\text{putC}, \text{putC}, \text{takeS, putC}] )</td>
</tr>
<tr>
<td>([\text{putC}, \text{putC}, \text{takeC}, \text{putC}] )</td>
<td>([\text{takeC}, \text{putC}, \text{takeC}, \text{takeC}] )</td>
</tr>
<tr>
<td>([\text{putC}, \text{putC}, \text{takeC}, \text{putS}] )</td>
<td>([\text{takeC}, \text{putC}, \text{takeS, takeC}] )</td>
</tr>
</tbody>
</table>

<table>
<thead>
<tr>
<th>Translations without potential overlaps</th>
</tr>
</thead>
<tbody>
<tr>
<td>([\text{putS, putC}, \text{takeC}, \text{putC}] )</td>
</tr>
<tr>
<td>([\text{takeC}, \text{putS, takeC}, \text{takeC}] )</td>
</tr>
<tr>
<td>([\text{putC}, \text{putC}, \text{takeC}, \text{putC}] )</td>
</tr>
<tr>
<td>([\text{putC}, \text{putC}, \text{takeC}, \text{putS}] )</td>
</tr>
</tbody>
</table>

Table 2. Non-rejected translations for 3 check-MVars and interprocess check restriction

**Theorem 5.6.** There is no valid \( \text{gstb} \)-translation with the interprocess check restriction for less than three check-MVars.

**Proof.** (Sketch) Our simulator refuted all translations for less than three check-MVars: All 72 possibilities of translations using 2 check-MVars could be refuted, where two processes suffice to refute all translation variants: \( x(y). \text{Stop} \) | \( x(y) \) and \( x(z)y(u). \text{Stop} \), where the first process is should-convergent and the latter is neither may- nor should-convergent.

A reason for the failure of translations with less than three check-MVars may be the following result:

**Theorem 5.7.** There is no executable, communicating, and overlap-free \( \text{gstb} \)-translation with the interprocess check restriction for \( n < 3 \).

**Proof.** For \( n = 1 \) we check the last four translations in Table 1. The first four are non-communicating. For \( ([\text{putS, takeC}], [\text{takeS, putC}]) \): after executing \( \text{putS, takeS} \) we can execute \( \text{putS} \) again. For \( ([\text{putS, takeC}], [\text{putC, takeS}]) \): after \( \text{putC, putS, takeC} \) we can execute \( \text{putC} \) again. For \( ([\text{takeC, putS}], [\text{takeS, putC}]) \): A deadlock occurs immediately. Thus the translation is not executable. For \( ([\text{takeC, putS}], [\text{putC, takeS}]) \): After executing \( \text{putC, takeC} \), we can execute \( \text{putC} \) again. For \( n = 2 \), we have checked all cases using our simulator, and no executable, communicating, and overlap-free translation for \( n = 2 \) is found. The numbers are: 18 translations are ruled out, since these are non-communicating, 21 lead to a deadlock, and 33 may lead to an overlap. This check could also be done by hand by scanning all 72 translations.
**Proposition 5.8.** The first four translations in Table 2 are executable, communicating, and overlap-free, whereas the next two translations are executable, communicating, but are overlapping.

*Proof.* The executability and overlap-freeness can be easily checked by simulating the possible executions. In any case, only if all 8 actions are finished, the next send or receive can start.

For the next two translations the overlap shows up after executing \( \text{putC}_1 \), \( \text{takeC}_1 \), since we can again execute \( \text{putC}_1 \).

In Appendix B.3 we sketch the arguments that the induced translation from the first gstb-translation from Table 2, i.e. \( \phi_{T_1} = (T_{\text{send}}, T_{\text{receive}}) = ([\text{putS}, \text{putC}_1, \text{takeC}_2, \text{putC}_3], [\text{takeC}_1, \text{putC}_2, \text{takeC}_3, \text{takeS}]) \), is indeed correct and leaves may- and should convergence invariant. The main help in reasoning is that there is no unintended interleaving of send and receive sequences according to Proposition 5.8. Application of Proposition 5.8 shows the following theorem.

**Theorem 5.9.** Translation \( \phi_{T_1} \) is adequate and on closed processes it is fully-abstract.

For 4 MVars there are 9600 translations and our simulation refutes 9299 and thus there remain 341 candidates for correct translations.

### 5.5 Dropping the Interprocess Check Restriction

If we drop the interprocess check restriction, but only allow one check-MVar without reuse, then we get 20 candidates for translations, which are all refuted by our simulation.

Allowing reuse of the single check-MVar seems not to help to construct a correct translation: We simulated this for up to 6 uses (which leads to 420420 candidates for a correct translation) – our simulation refutes all of them.

We conjecture that there is no correct translation for one check-MVar where re-uses are permitted, and where the interprocess check restriction is dropped.

For two MVars and one use but without the interprocess check-MVar there are 420 translations. Our simulation refutes all but two translations, which are shown in Table 3. In the first translation the second check-MVar is used as a

<table>
<thead>
<tr>
<th>Sender</th>
<th>Receiver</th>
</tr>
</thead>
<tbody>
<tr>
<td>[\text{putC}_1, \text{putS}, \text{takeC}_2, \text{takeC}_1]</td>
<td>[\text{takeS}, \text{putC}_2]</td>
</tr>
<tr>
<td>[\text{takeC}_1, \text{putS}]</td>
<td>[\text{putC}_2, \text{putC}_1, \text{takeS}, \text{takeC}_2]</td>
</tr>
</tbody>
</table>

**Table 3.** Non-refuted translations for two check-MVars, no interprocess check restriction
mutex for the senders, ensuring that only one sender can operate at a time, while the second translation does the same on the receiver side.

**Proposition 5.10.** The translations in Table 3 are executable, communicating, and overlap-free.

**Proof.** It is easy to verify that the translations are executable and communicating. For the first translation, assume that \( \text{putC}^1, \text{putS} \) and \( \text{takeS} \) are performed in this order (no other order is possible). Then an additional sender cannot execute its first command before the original sender performs \( \text{takeC}^1 \) and this again is only possible after the receiver program is finished. An additional receiver can only be executed after a \( \text{putS} \) is performed, which cannot be done by the current sender and receivers.

For the second translation, assume that \( \text{putC}^2, \text{putC}^1 \) and \( \text{takeC}^1 \) are performed in this order (no other order is possible). An additional receiver can only start after \( \text{takeC}^2 \) was executed by the original receiver, which can only occur after the original sender and receiver program are fully evaluated. An additional sender can only start after \( \text{putC}^1 \) has been executed again, but the current sender and receiver cannot execute this command.

In the appendix we sketch in Theorem B.4 that the induced translation from the first \( \text{gstb-translation} \), i.e. the translation \( \phi_{\mathcal{I}_2} \) with

\[
\mathcal{I}_2 = ([\text{putC}^1, \text{putS}, \text{takeC}^2, \text{takeC}^1], [\text{takeS}, \text{putC}^2])
\]

is (closed) convergence equivalent.

Applying Proposition 5.3 this shows

**Theorem 5.11.** Translation \( \phi_{\mathcal{I}_2} \) is adequate and on closed processes it is fully-abstract.

We are convinced that the same holds for the second translation. We conclude the statistics of our search for translations without the interprocess restriction:

- For 3 MVars there are 10080 translations and 9992 are refuted, i.e. 98 translations are potentially correct. Among the 98 translations, there is the translation \( ([\text{putC}^1, \text{putS}, \text{takeC}^2, \text{takeC}^1], [\text{takeS}, \text{putC}^2]) \) which is quite intuitive: check-MVar 1 is used as a mutex for all senders on the same channel, check-MVar 3 is used as a mutex for all receivers, and check-MVar 2 is used to acknowledge that the message was received.
- For 4 MVars there are 277200 translations and 273210 are refuted, i.e. 3990 translations are potentially correct.

6 Discussion and Conclusion

We have investigated the problem of translating the \( \pi \)-calculus into the \( CH \)-calculus modeling Concurrent Haskell and proved that there is a correct translation \( \tau_0 \) from the \( \pi \)-calculus into \( CH \), using private names for the storage of
every translated communication construct. This translation is an embedding for closed processes w.r.t. $\sim_c$, and preserves the may-convergence behavior as well as the should-convergence behavior. The translation $\tau_0$ meets the gold standard, since even for open processes, it is adequate w.r.t. the respective contextual equivalences.

For translations that use global names, we started an investigation on exhibiting (potentially) correct translations. The result is that in the design space of all (global) translations we identified several minimal translations that are potentially correct. We characterized all incorrect “small” translations, where we left the issue open whether there is a correct translation with only one check-MVar with unrestricted usage.

For two particular global translations, we have shown that they are convergence equivalent, i.e. they preserve and reflect the may-convergence behavior and the should-convergence behavior on closed processes. We also proved adequacy of the translations (on open processes) and full-abstraction if closed processes are translated. The exact form of the translations were found by our implemented tool to search for translations and to refute their correctness. The tool also showed that there is no correct gstb-translation with the interprocess check restriction for less than 3 check-MVars.

Further work may be to consider extended variants of the $\pi$-calculus. We are convinced that adding recursion and sums can easily be built into the translation, while it might be challenging to encode mixed sums or (name) matching operators. For name matching operators, our current translation would require to test usual bindings in $CH$ for equality which is not available in core-Haskell. Solutions may either use an adapted translation or an adapted core language that supports so-called observable sharing [5,7]. The translation of mixed-sums into $CH$ appears to require more complex translations, where the send and receive-parts are not linear lists of actions.

Another interesting research question for future work is whether the pi-calculus can be embedded in call-by-need lambda calculi with $amb$. Note that [4] show that the $amb$-operator is encodable in the $\pi$-calculus.

Acknowledgments

We thank an anonymous reviewer for advises to improve the translation. In particular for providing the counter-example in the last row of Table 1.
References

\[\sigma_0(P) = vstop.(\text{putMVar stop} () | \text{stop m} () | \sigma(P))\]
\[\sigma(vx.P) = vx, \text{chanx.}(\text{chanx m} - Ix = \text{Channel chanx} | \sigma(P))\]
\[\sigma(\overline{xy}.P) = \tau(\overline{xy}.P)\]
\[\sigma((0) = \ll return ()\]
\[\sigma(\text{Stop}) = \ll \text{takeMVar stop}\]
\[\sigma(P | Q) = \sigma(P) | \sigma(Q)\]
\[\sigma(!P) = \nu f.(\ll f | f = \text{do } \{\text{forkIO } \tau(P); f\})\]

Fig. 7. Translations \(\sigma\) and \(\sigma_0\)

A Convergence Properties of the Translation \(\tau_0\)

A.1 A Top-Down Variant \(\sigma\) of \(\tau\)

We define a variant \(\sigma\) of the translation \(\tau\). It is intended to simplify the proofs. Instead of generating \(CH\)-programs that during their run generate concurrent processes and MVars (as \(\tau\) does), the translation \(\sigma\) generates \(CH\)-processes, which is closer to a direct implementation. The translation \(\sigma\) only modifies a top-part, and refers to \(\tau\) for the deeper translation parts that generate expressions. The main reason for introducing the translation \(\sigma\) is that it simplifies the proofs of correctness (of translations \(\tau\) and \(\tau_0\)).

Definition A.1. The translation \(\sigma_0\) is similar to \(\tau_0\), and the translation \(\sigma\) also generates processes. Both translations are defined in Fig. 7. The translations of contexts are defined analogously where the context hole is translated into the context hole, i.e. \(\sigma([\cdot]) = [\cdot]\).

Definition A.2. The context \(C_{\sigma_0}^\sigma\) for translation \(\sigma_0\) is defined as

\[C_{\sigma_0}^\sigma = \nu f, \text{stop}.(\ll \text{putMVar stop} () | \text{stop m} () | \ll [\cdot]).\]

The following strong relationship between the translations \(\tau_0\) and \(\sigma_0\) (\(\tau\) and \(\sigma\), resp.) holds and is helpful in proofs for the translation \(\tau\):

Lemma A.3. For all \(P \in \Pi_{\text{Stop i)} \tau_0(P) \stackrel{sr,\ast}{\rightarrow} \sigma_0(P)\) and ii) \(\tau(P) \stackrel{sr,\ast}{\rightarrow} \sigma(P)\).

Proof. For the first part, it suffices to verify that \(\tau_0(P)\) is a process that reduces to \(\sigma_0(P)\) as follows:

\[\tau_0(P) = \ll \text{do } \{\text{stop } \leftarrow \text{newMVar } (); \text{forkIO } \tau(P); \text{putMVar stop } ()\}\]
\[\ll \text{vsstop.}\text{stop m} () | \ll \text{do } \{\text{forkIO } \tau(P); \text{putMVar stop } ()\}\]
\[\ll \text{vsfork.}\text{stop m} () | \ll \text{do } \{\text{forkIO } \tau(P); \text{putMVar stop } ()\}\]
\[\ll \text{vsstop.}\text{stop m} () | \ll \text{do } \{\text{forkIO } \tau(P); \text{putMVar stop } ()\}\]

The \(\stackrel{sr,\ast}{\rightarrow}\)-sequence is derived from the second part of the lemma, and by plugging-in the reduction sequence in the larger context. It remains to show
the second part. We show this by induction on the size of \( P \) and by checking all the cases. If \( P \) starts with an input or an output prefix or is \( \texttt{Stop} \), then \( \langle \tau(P) \rangle = \sigma(P) \), and thus the claim holds. If \( P \) is the process \( 0 \), then \( \langle \tau(P) \rangle = \langle \texttt{return} () \rangle = \sigma(P) \). If \( P \) is a parallel composition \( P_1 \parallel P_2 \), then

\[
\langle \tau(P_1 | P_2) \rangle = \langle \texttt{do} \{ \texttt{forkIO} \ \tau(P_2); \ \tau(P_1) \} \rangle
\]

The final \( \texttt{sr}* \)-sequence obtained as follows: by the induction hypothesis we have \( \langle \tau(P_i) \rangle = \sigma(P_i) \) for \( i = 1, 2 \) by combing and processing them sequentially we get \( \langle \tau(P_2) \rangle = \langle \tau(P_1) \rangle = \sigma(P_1) \). If the process starts with \( \nu x \), then

\[
\langle \tau(\nu x. P) \rangle = \langle \texttt{do} \{ \texttt{chan} x \leftarrow \texttt{newMVar} \bot; \texttt{takeMVar} \ \texttt{chan} x; \\
\texttt{letrec} \ x = \texttt{Channel} \ \texttt{chan} \ x \texttt{in} \ \tau(P) \} \rangle
\]

This finishes the induction proof. \( \square \)

Lemma \[\text{A.3}\] and correctness of (fork), (nmvar), (dtmvar), (beta), (cpce), and (mkbinds) (see Proposition \[3.6\]) imply:

**Proposition A.4.** For all \( P \in \Pi_{\text{Stop}} \): 1. \( \sigma(P) \sim_\epsilon \tau(P) \) and \( \langle \sigma_0(P) \rangle \sim_\epsilon \tau_0(P) \). Hence also \( \sigma(P) \sim_{c, \tau_0} \tau(P) \) and \( \sigma_0(P) \sim_{c, \tau_0} \tau_0(P) \).

### A.2 Preservation of May-Convergence

The goal of this section is to show that may-convergence of a closed \( \Pi_{\text{Stop}} \)-process \( P \) implies may-convergence of the translated process \( \tau_0(P) \). The main part is to show that for a single step \( P \ \texttt{dia} \ P' \) or \( P \ \texttt{dec} \ P' \) for closed \( \Pi_{\text{Stop}} \)-processes \( P, P' \), there are corresponding reduction steps of \( \sigma_0(P) \) in the CH-calculus.
Lemma A.5. Let \( P \in \Pi_{\text{Stop}} \) be a closed process with \( P \xrightarrow{\text{dis}} P' \). Then there is a sequence \( \sigma_0(P) \xrightarrow{sr} \sigma_0(P') \mid G \) in the CH-calculus, where \( G \) is a closed process that can be seen as garbage, i.e. \( \sigma_0(P') \mid G \xrightarrow{gc} \sigma_0(P') \).

Proof. Let \( P = \mathbb{D}[x(y).P, \mid xz.P] \) and \( P' = \mathbb{D}[P,z/y \mid P_s] \). Since \( P \) is closed, there is a binder \( \nu x \) in \( \mathbb{D} \), i.e. we can assume that \( \mathbb{D} = \mathbb{D}_1[\nu x. \mathbb{D}_2[\nu x]] \) for some \( \mathbb{D}_x \)-contexts \( \mathbb{D}_1, \mathbb{D}_2 \). Since \( \mathbb{D}_x \)-contexts have no input- or output-prefix on the hole-path, this shows:

\[
\sigma_0(P) = C'\sigma_0[D_1[\nu x, \text{chanx.chanx } m - \mid x = \text{Channel chanx} \\
\mid D_2[\mid \text{do \{ checkx } \leftarrow \text{newMVar } () \right. \\
\text{putMVar } \text{(unchan } x \text{)} (z, \text{checkx}); \\
\text{putMVar } \text{checkx } (); \tau(P_s)) \mid \text{do \{ } \text{\tau}(P_s) \mid \text{do \{ } \text{\tau}(P_s)[z/y]\})]]
\]

where \( D_1, D_2 \) are the \( \sigma \)-translations of \( D_1, D_2 \). Inspecting \( \sigma \) shows, that \( D_1, D_2 \) are \( \mathbb{D} \)-contexts. There is a sequence of \( \xrightarrow{sr} \)-reductions (see Fig. 8) such that

\[
\sigma_0(P) \xrightarrow{sr,15} Q = \sigma_0[D_1[\nu x, \text{chanx.chanx } m - \mid x = \text{Channel chanx} \\
\mid D_2[\mid \text{do \{ } \text{\tau}(P_s) \mid \text{do \{ } \text{\tau}(P_s)[z/y]\})]]
\]

where \( G = (\nu \text{checkx.chanx } m (\)) \) is garbage. Since

\[
\sigma_0(P') = C'\sigma_0[D_1[\nu x, \text{chanx.chanx } m - \mid x = \text{Channel chanx} \\
\mid D_2[\mid \text{do \{ } \text{\tau}(P_s) \mid \text{do \{ } \text{\tau}(P_s)[z/y]\})]]
\]

Lemma [A.3] implies \( Q \xrightarrow{sr,} \sigma_0(P') \mid \nu \text{checkx.chanx } m (\) \xrightarrow{gc} \sigma_0(P') \).

Lemma A.6. Let \( P, P' \in \Pi_{\text{Stop}} \) be closed, such that \( P \xrightarrow{\text{dis}} P' \). Then \( \sigma_0(P) \xrightarrow{sr,*} \sigma_0(P') \), if rule (replunfold) is used, and \( \sigma_0(P) \equiv \sigma_0(P') \), otherwise. In particular, \( \sigma_0(P) \sim_c \sigma_0(P') \), and hence \( \sigma_0(P) \sim_{c,0} \sigma_0(P') \).

Proof. For rules (assocl) and (assocr), we have

\[
\sigma_0(\mathbb{D}(P_1 \mid P_2 \mid P_3)) = C[\sigma(P_1) \mid \sigma(P_2)] \mid \sigma(P_3) = C[\sigma(P_1) \mid \sigma(P_2)] \mid \sigma(P_3) = \sigma_0(\mathbb{D}(P_1 \mid P_2 \mid P_3))
\]

for some \( CH \) \( \mathbb{D} \)-context \( C \) and thus, \( \sigma_0(P) \equiv \sigma_0(P') \). For (nuup1), we have

\[
\sigma_0(\mathbb{D}(\nu z.P_1 \mid P_2)) = C[\sigma(\nu z.P_1) \mid \sigma(P_2)]
\]

\[
= \sigma_0(\mathbb{D}(\nu z, \text{chanz.(chanz } m - \mid z = \text{Channel chanz } \mid \sigma(P_1) \mid \sigma(P_2)))
\]

\[
\equiv \sigma_0(\mathbb{D}(\nu z, \text{chanz.chanx } m - \mid z = \text{Channel chanx} \\
\mid \nu z, \text{chanx.chanx } m - \mid z = \text{Channel chanx } \mid \sigma(P_1)))
\]

for some \( CH \) \( \mathbb{D} \)-context \( C \). Thus, \( \sigma_0(P) \equiv \sigma_0(P') \). For (nuup2), we have

\[
\sigma_0(\mathbb{D}(\nu x.\nu z.P_1)) = C[\sigma(\nu x.\nu z.P_1)]
\]

\[
= C[\nu x, \text{chanx.chanx } m - \mid x = \text{Channel chanx} \\
\mid \nu z, \text{chanz.(chanz } m - \mid z = \text{Channel chanz } \mid \sigma(P_1)))
\]

\[
\equiv \sigma_0(\mathbb{D}(\nu x.\nu z.P_1))
\]
\[ \sigma_0 (P) = C^{\sigma}_{\text{out}} [D_x[vx, \text{chan.x.chan}] \mid x = \text{Channel chanx} \\
| D_2 \leftarrow \text{do } \{ \text{checkx } \leftarrow \text{newMVar }(); \text{putMVar } (\text{unchan } x) (z, \text{checkx}); \text{putMVar } \text{checkx }(); \tau (P_x) \} \]

\[ C^{\sigma}_{\text{out}} [D_x[vx, \text{chan.x.chan}] \mid x = \text{Channel chanx} \\
| D_2 \leftarrow \text{do } \{ \text{putMVar } (\text{unchan } x) (z, \text{checkx}); \text{putMVar } \text{checkx }(); \tau (P_y) \} \]

\[ \equiv \]

\[ \equiv \]

\[ Q \]

**Fig. 8.** Reduction sequence \( \sigma_0 (P) \xrightarrow{sr,15} Q \) which is used in Lemma A.5.
for some \( CH \) \( \mathbb{D} \)-context \( C \). Thus, we have \( \sigma_0(P) \equiv \sigma_0(P') \). For rule (commute), we have \( \sigma_0(P) \equiv \sigma_0(P') \), since
\[
\sigma_0(\mathbb{D}[P_1 \mid P_2]) = C[(\sigma(P_1) \mid \sigma(P_2))] \equiv C[(\sigma(P_2) \mid \sigma(P_1))] = \sigma_0(\mathbb{D}[P_2 \mid P_1])
\]
for some \( \mathbb{D} \)-context \( C \). Thus the claim holds.

For the rule (replunfold), we have:
\[
\sigma_0(\mathbb{D}[!P]) = C[\nu f.(\leftarrow f \mid f = \text{do} \{\text{fork} \text{IO} \tau(P)\}; f)]
\]
\[
\xrightarrow{sr,\text{cpce}} C[\nu f.(\leftarrow \text{do} \{\text{fork} \text{IO} \tau(P)\}; f) \mid \text{do} \{\text{fork} \text{IO} \tau(P)\}; f)]
\]
\[
\xrightarrow{sr,\text{fork}} C[\nu f.(\leftarrow \tau(P) \mid f \mid f = \text{do} \{\text{fork} \text{IO} \tau(P)\}; f)]
\]
\[
\equiv C[\nu f.(\leftarrow \tau(P) \mid f \mid f = \text{do} \{\text{fork} \text{IO} \tau(P)\}; f)]
\]
\[
\xrightarrow{sr,\text{beta}} C[\tau(P) \mid f \mid f = \text{do} \{\text{fork} \text{IO} \tau(P)\}; f)] = \sigma_0(\mathbb{D}[P ! ! P])
\]

where \( f \not\in FV(P) \). This shows \( \sigma_0(\mathbb{D}[!P]) \xrightarrow{sr,*} \sigma_0(\mathbb{D}[P ! ! P]) \), using the previous items, and Lemma A.3 for the last \( sr,* \)-transformation. The equivalence \( \sigma_0(P) \sim_c \sigma_0(P') \) follows from correctness of the used reduction steps (see Proposition A.6).

**Lemma A.7.** Let \( P \) be a closed \( CH \)-process such that \( P \xrightarrow{(sr,* \cdot \sim_c)^n} Q \). Then the following implications hold: i) If \( Q \downarrow \), then \( P \downarrow \) and ii) If \( Q \uparrow \), then \( P \uparrow \).

**Proof.** We only show the first part. The proof for the second part can be obtained by replacing \( \downarrow \) by \( \uparrow \). The proof is by induction on the length of the relational sequence. Let \( P \xrightarrow{(sr,* \cdot \sim_c)^n} P_1 \xrightarrow{sr,*} P_2 \sim_c Q \), where \( Q \downarrow \). From \( P_2 \sim_c Q \) we derive that also \( P_2 \downarrow \), and hence also \( P_1 \downarrow \), since \( P_1 \xrightarrow{sr,*} P_2 \).

**Lemma A.8.** If a closed process \( P \) of the calculus \( \Pi_{\text{stop}} \) is successful, then \( \sigma_0(P) \xrightarrow{sr,*} Q \) where \( Q \) is successful.

**Proof.** Process \( P \) must contain \texttt{Stop} in a \( \mathbb{D} \)-context. Thus the translation generates a thread in a \( \mathbb{D} \)-context, that executes \texttt{takeMVar} \texttt{stop} and \( Q \xrightarrow{sr,\text{tmvar},sr,\text{pmvar}} D(\texttt{return}()) \). For the second part of the lemma, it suffices to observe that \( \sigma \) and \( \tau \) do not generate a \texttt{putMVar} for the MVar \texttt{stop} (except in the context \( C_{\text{out}}^\sigma \)) and thus, the MVar can always be emptied and the main-thread thereafter can perform the \texttt{putMVar}-command to become successful.

**Proposition A.9.** Implication of may-convergence: Let \( P \) be a closed \( \Pi_{\text{stop}} \)-process. If \( P \downarrow \) then \( \tau(P)\downarrow_0 \).

**Proof.** Lemmas A.5 to A.8 imply that \( \sigma_0(P) \xrightarrow{sr,*} Q \), where \( Q \) is successful, using induction on the length of a standard reduction of \( P \), hence \( \sigma(P)\downarrow_0 \). By Proposition A.4 this also shows \( \tau_0(P)\downarrow_0 \), which is the same as \( \tau(P)\downarrow_0 \).

Note that in the previous proof it is irrelevant whether we remove the garbage that stems from translated \( \xrightarrow{dia} \)-steps or not: In the former case we know that \( (gc) \) is a correct program transformation and thus does not change convergence, and in the latter case we carry the garbage with the whole reduction sequence.
Fig. 9. Proving reflection of may-convergence: structure of the induction step

**A.3 The Translation Reflects May-Convergence**

In this section we show that may-convergence of $\sigma_0(P)$ for a closed $\Pi_{\text{Stop}}$-process $P$ implies may-convergence of $P$. We show that, by rearranging and extending the standard reduction sequence to a successful process, a standard reduction sequence of $\sigma_0(P)$ is obtained that corresponds to a standard reduction sequence of $P$ in the $\pi$-calculus. There are three essentially different actions that are executed by the standard reduction sequence of $\sigma_0(P)$, where the single reduction steps may be distributed in the reduction sequence: interaction-reductions, replication of the $!P$-operator, and reducing $\tau$-images to $\sigma$-images.

The following properties are easily checked for a translated closed $\Pi_{\text{Stop}}$-process $P$:

- $\tau_0(P)$ is closed;
- $\tau_0(P)$ contains the variable $\text{stop}$ only in expressions of the form $\leftarrow \text{takeMVar} \text{stop}$, and
- $\tau_0(P)$ is well-formed and well-typed.

**Proposition A.10.** Let $P \in \Pi_{\text{Stop}}$ be closed, $n \in \mathbb{N}$ and $\sigma_0(P) \xrightarrow{sr,n} Q_1$ such that $Q_1$ is successful. Then there is a standard reduction sequence $P \xrightarrow{sr,*} P_1$ and another standard reduction sequence $\sigma_0(P) \xrightarrow{sr,*} Q'_1$ such that $Q'_1$ and $P_1$ are successful and $Q'_1 = \sigma_0(P_1)$.

**Proof.** We show that the following diagram holds by induction on the number of all reduction steps in $\text{Red} = \sigma_0(P) \xrightarrow{sr,n} Q_1$.

In this diagram and in the diagram in Fig. 9 a plain arrow means a given and a dashed arrow means an existing reduction.

The base case is that $\sigma_0(P)$ is of the form $C_{\text{out}}[D][\leftarrow \text{takeMVar} \text{stop}]$, and it standard-reduces in two steps to a successful process. In this case, the only possible reason is that $P$ contains $\text{Stop}$ in a $\text{PCtxt}_\pi$-context, and $\sigma_0$ maps it to this subprocess. Then $P$ is successful.

Now we show the induction step. A picture of the proof structure is shown in Fig. 9. By the induction hypothesis, we have $P_1 \xrightarrow{sr,*}$, but we have to calculate the upper square for all possibilities of actions (interaction reductions, replication,
τ-to-σ-reduction), where we rearrange the CH-standard reduction sequence, extend it with missing reduction steps, which do not change success, and construct a corresponding standard reduction sequence in the π-calculus.

Abstractly, a single induction step is intended to do the following: first we identify reduction steps that make a complete or a partial execution that performs one of the three actions. There may be two threads affected by interaction, and otherwise only one thread is affected. We have to identify one particular (distributed) reduction subsequence $S$ that can be executed as a prefix; i.e. $\text{Red}$ can be rearranged to $S; \text{Red}'$ having the same total effect. There are two possibilities: Either there is such a full subsequence $S$ of reduction steps, or we only identify a prefix of such a subsequence, and the corresponding threads do not have further reduction steps (of the action) until $Q_1$ is reached. In the first case, everything is fine, and a process $P_1$ can be determined. In the second case, the subsequence has to be removed, or extended by the missing reduction steps. Since the thread(s) is/are stopped, these additional reduction steps are independent of all other reduction steps. Hence these can be virtually shifted after $Q_1$ (perhaps turning into non-sr-reductions) and determine $Q'_1$. We will add the extra reduction steps to the prefix of the reduction sequence, and then obtain a complete subsequence.

Since the intention of the proof is to construct a reduction sequence between images of translation $\sigma$, this enforces that sometimes $\tau$-translated parts have to be sr-reduced to their $\sigma$-translation, and that a (gc)-step has to be inserted as a $\sim_c$-step.

Referring to Definition A.1 of translation $\sigma$, there are several cases. Let us consider the case that an image of an ia-reduction is the first action. Analyzing $\text{Red}$ shows that, ideally, if we only look for the (newmvar)-, (tmvar)-, and (pmvar)-reductions, the reduction sequence starts as follows

\begin{align*}
(\text{ia-0}) & \quad \text{checkx} \leftarrow \text{newMVar}(); \\
(\text{ia-1}) & \quad \text{putMVar} (\text{unchan x}) (y, \text{checkx}) \text{ in thread a;} \\
(\text{ia-2}) & \quad \text{takeMVar} (\text{unchan x}) \text{ in thread b (}\neq a); \\
(\text{ia-3}) & \quad \text{takeMVar} \text{ checkx in thread b;} \\
(\text{ia-4}) & \quad \text{putMVar} \text{ checkx} () \text{ in thread a.}
\end{align*}

However, there may be deviations, like interleaving of reduction steps in parallel threads, or an incomplete subsequence. We analyse the possibilities in a reduction sequence, where we omit the non-interfering interleaved actions. Of course, we assume that the programs have the commands for (ia-0);(ia-1);(ia-4) as a sequence in one thread-program and the commands for (ia-2);(ia-3) as a sequence in the other thread-program.

1. (ia-0) is the only action for checkx; no further action; thread $a$ is blocked.
2. (ia-0);(ia-1) is performed in thread $a$; no other thread uses the MVar (unchan x); thread $a$ is blocked.
3. (ia-0);(ia-1);(ia-2) is performed first in thread $a$, and some other thread $b$ won the race for (ia-2), and then no further progress; threads $a, b$ are blocked. takeMVar checkx is not possible in another thread, since the name is unknown to other threads. However, the MVar (unchan x) is already available for other threads.
4. \((ia-0);(ia-1);(ia-2);(ia-3)\) is performed with \((ia-0);(ia-1)\) in thread \(a\), and then \((ia-2);(ia-3)\) in thread \(b\); thread \(a\) is blocked, but thread \(b\) can proceed. MVar \(\text{checkx}\) is not garbage.

5. \((ia-0);(ia-1);(ia-2);(ia-3);(ia-4)\) is performed with \((ia-0);(ia-1);(ia-4)\) in thread \(a\), and \((ia-2);(ia-3)\) in thread \(b\). Thread \(a\) is released after the full sequence, thread \(b\) after \((ia-3)\). Moreover, after \((ia-4)\), the MVar \(\text{checkx}\) can no longer be used and is garbage.

We show that modifying the reduction sequence \(\text{Red}\), keeping the sr-property, results in another reduction sequence \(\text{Red'}\) that starts with the ideal 5 reductions, followed by \(\text{Red''}\) also ends in a successful process, such that \(\text{Red''}\) contains less standard-reduction steps.

Now we focus the \(\text{ia}\)-related reduction steps in \(\text{Red}\). We remove the following subsequences: In case there is a thread such that after an \((ia-0)\) or after an \((ia-0);(ia-1)\) step, then the thread is blocked; Then we remove these reduction steps from \(\text{Red}\). We do this in all cases.

Now we focus the earliest occurrence in \(\text{Red}\) of an \((ia-2)\)-related reduction step. If the \((ia)\)-subsequence corresponding to the first one is complete, then we move it to the start of the sequence. Of course the non-MVar related sr-steps in the corresponding threads \(a,b\) are also moved. In the following we do not mention these sr-reductions.

Now let us consider the incomplete ones:

- If exactly \((ia-0);(ia-1);(ia-2)\) are in \(\text{Red}\), then thread \(a,b\) do not further contribute to the success, however, MVar \(\text{unchan x}\) may be used for the success (triggered by another thread). We add the reductions \((ia-3)\) and \((ia-4)\) in order to complete the ia-sequence, which makes the reduction sequence longer, but does not add \((ia-2)\)-reduction steps.
- If exactly \((ia-0);(ia-1);(ia-2);(ia-3)\) are in \(\text{Red}\), then we add the missing \((ia-4)\) step without changing success and the rest of the reduction sequence.

In order to apply the induction hypothesis, we insert a \((\text{gc})\)-step to remove the check-MVar that was only used for one interaction step in the translation. This \((\text{gc})\) is correct, hence we can represent it by a \(\sim\)-step.

Let \(\leftarrow \tau(P_a)\) and \(\leftarrow \tau(P_b)\) be the processes that are left by the four reduction steps that simulate the \((ia)\)-reduction. It remains to check whether \(\leftarrow \tau(P_a)\) and \(\leftarrow \tau(P_b)\) reduce to \(\sigma(P_a)\) and \(\sigma(P_b)\), respectively. Here we use the same construction as above: The reduction steps that are already in the sequence are shifted to the left. Since the threads are blocked if there are missing reductions, we can add the missing reduction steps to the reduction sequence, keeping the property that the resulting process of the whole sequence is successful. For details, we refer to Lemma \(\text{A.3}\). We obtain the upper square of the diagram, where also the number of reductions is strictly smaller for the remaining reduction sequence.

We now consider the case that the reduction is an image of a \((\text{replunfold})\) step. Then the corresponding reduction is the \((\text{fork})\)-reduction. In fact, it is a sequence \(\text{opc} \cdot \text{fork} \cdot \beta\). We can, following the reduction pattern in the proof of Lemma \(\text{A.6}\) move the reduction steps to the front, which are corresponding to
a single replication. If reduction steps are missing, we can add the missing steps without disturbing the final success. Also, we can add the necessary reductions that may be missing to turn the \( \Leftarrow \tau(P_r) \) into \( \sigma(P_r) \). Again we can construct the square diagram, as requested, where also in this case, the process that represents the final success may have changed.

Finally, we can apply the induction hypothesis, since the combined measure is strictly reduced, and we thus obtain a standard reduction in \( \Pi_{\text{Stop}} \) to a successful process.

Propositions \[A.9\] and \[A.10\] imply:

**Proposition 4.5.** Let \( P \in \Pi_{\text{Stop}} \) be closed. Then \( \tau_0 \) is convergence-equivalent for \( \Downarrow \), i.e. \( P \Downarrow \) is equivalent to \( \tau_0(P) \Downarrow \). This also implies that \( P \Uparrow \) is equivalent to \( \tau_0(P) \Uparrow \).

### A.4 Equivalence of Should-Convergence of the Translation

We argue that the translation \( \tau \) is invariant w.r.t. should-convergence, where we work with may-divergence and where Proposition \[4.5\] is very useful since it is the induction base.

**Proposition A.11 (Preservation of May-Divergence).** Let \( P \in \Pi_{\text{Stop}} \) be closed, \( n \in \mathbb{N} \) and \( \sigma_0(P) \stackrel{\text{sr},n}{\rightarrow} Q \) such that \( Q \Uparrow \). Then there is a standard reduction sequence \( \sigma_0(P) \stackrel{\text{sr},*}{\rightarrow} Q \) with \( Q \Uparrow \).

**Proof.** We use induction on \( n \) to show that \( P \stackrel{\text{sr},n}{\rightarrow} P_1 \) where \( P_1 \Uparrow \) implies \( \sigma_0(P) \stackrel{\text{sr},*}{\rightarrow} Q' \) where \( Q' \) is must-divergent. The base case \( n = 0 \) is covered by Proposition \[4.5\]. For the induction step, Lemmas \[A.5\] and \[A.6\] can be applied for the first reduction step and then the induction hypothesis can be applied to show the claim. Finally, we apply item ii) of Lemma \[A.7\] to derive the claim.

The last case is to show that reflection of may-divergence holds, i.e. \( \sigma_0(P) \Uparrow \implies P \Uparrow \). This is almost similar to the arguments for \( \sigma_0(P) \Downarrow \implies P \Downarrow \) where some more arguments are needed to show that the final process remains must-divergent after the rearrangements and additions to the reduction sequence.

**Proposition A.12.** Let \( P \in \Pi_{\text{Stop}} \) be closed, \( n \in \mathbb{N} \) and \( \sigma_0(P) \stackrel{\text{sr},n}{\rightarrow} Q_1 \) such that \( Q_1 \Uparrow \). Then there is a standard reduction sequence \( P \stackrel{\text{sr},*}{\rightarrow} P_1 \) with \( P_1 \Uparrow \), and another standard reduction sequence \( \sigma_0(P) \stackrel{\text{sr},*}{\rightarrow} Q'_1 \) such that \( Q'_1 \Uparrow \) and \( Q'_1 = \sigma_0(P'_1) \).

**Proof.** The following diagram gives an orientation on the goals of the proof and the induction on \( n \).

\[
\begin{array}{c}
P \\ \text{sr},* \\ \Uparrow \end{array} \quad \quad \begin{array}{c}
\sigma_0 \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad\]
Most arguments for a construction of the reduction sequence $\sigma_0(P) \xrightarrow{srt_{\cup \sim c}} \sigma_0(P_1)$ are already in the proof of Proposition [A.10]. However, the treatment of the incomplete (ia)-reduction sequences is different, so we make it explicit.

The cases are:

1. In the case that only the (ia-0) is present, we can omit it, since it is correct.
2. If at least (ia-0);(ia-1);(ia-2) is there, then we add the missing reductions.
   The argument is: If $\sigma_0(P) \xrightarrow{srt_{\sim c}} Q \uparrow$ and $Q \xrightarrow{srt_{\cup \sim c}} Q'$ then also $Q' \uparrow$, and we can perform the extra reductions (like (ia-3);(ia-4) from $Q$, and then shift them closer to the (ia-0);(ia-1);(ia-2)-sequence. This makes the ia-sequence complete, and does not change the induction measure.
3. The case that the incomplete sequence is (ia-0);(ia-1), and (ia-2) is not present, needs more arguments: There are two possibilities:
   (a) There is a reduction sequence starting from $Q$, which contains the needed (ia-2). Then we will shift first the (ia-2) such that $Q \xrightarrow{(ia-2)} Q'$, where obviously $Q' \uparrow$. Then we shift the (ia-2) further to the left, such that we have the sequence (ia-0);(ia-1);(ia-2), and we can then use the other cases. Here we have modify the induction measure: first the number of (ia-2) in the reduction sequence to $Q$, and as second component the total number of (ia)-reductions in the sequence $\sigma_0(P) \xrightarrow{srt_{\sim c}} Q$. We have removed more than one (ia)-reduction in the induction, hence we can apply the induction hypothesis.
   (b) There is no reduction sequence starting from $Q$ with the needed (ia-2). In this case the argument uses that the current CH-program is a translation of a $\Pi_{\text{Stop}}$-process. We remove the (ia-1) from the reduction sequence. The reduction leads to $Q_1$, which is the result of the reduction sequence leading to $Q$, but without the (ia-1)-reduction. Now there may be another thread that can access the MVar ($\text{unchan } x$), again using (ia-0);(ia-1) in the other thread. However, since from $Q$ there is no (ia-2) possible for MVar ($\text{unchan } x$), the same must hold for $Q_1$. Hence, the removal permits only small local changes, but there is still no successful reduction from $Q_1$. Hence $Q_1 \uparrow$. Of course, we can now also remove the (ia-0)-reduction.

Note that we also have to use the (gc)-reduction removing the check-MVar in case of a complete (ia)reduction in CH, which adds a $\sim c$ to the combined reduction sequence.

Using these arguments the proof is otherwise completely analogous to the proof of Proposition [A.10] and we are done.

**Proposition [4.6].** Let $P \in \Pi_{\text{Stop}}$ be closed. Then $\tau_0$ is convergence-equivalent for $\downarrow$, i.e. $P \downarrow$ is equivalent to $\tau_0(P) \downarrow$. 

B Proofs and Details on Translations with Global Names

B.1 Failing Translations with Less than Three Check-MVars and the Interprocess Check Restriction: Examples

Before we prove that the first translation of Table 2 is invariant for may- and should-convergence, we show, how similar translations with less check-MVars fail. The reason is usually that the translated sequences for send and receive are not forced to be completely executed. For example, the translation of \( \pi_y \mid x(y).\pi_y.\text{Stop} \) with one check-MVar where the sender program is \([\text{putS}, \text{takeC}]\); and the receiver program is \([\text{putC}, \text{takeS}]\) confuses the sequence for the program consisting of two threads, and where the \( \pi \)-process is not may-convergent, but the translated program may-converges. The program is written in the first line as two threads, and the sequence of execution in the second line:

\[
\begin{array}{|c|c|c|c|c|c|}
\hline
\text{send} & \text{receive} \\
\hline
[\text{putS}, \text{takeC}] & [\text{putC}, \text{takeS}, \text{putS}, \text{takeC}, \text{Stop}] \\
\hline
1 & 2 & 3 & 4 & 5 & 6
\end{array}
\]

The situation is similar for all other translations using one or two check-MVars, and where every check-MVar is written and read only once per translated interact. Either the communication is not synchronous, or the sequences of send and receive that do not belong to the intended interact can be interleaved.

For documentation purposes, we detail an example for a bad behavior for a translation using two check-MVars and which is similar to the first one, where this example is a result of using our simulator. The translation is

\[
[\text{putS, putC}^1, \text{takeC}^2] \mid [\text{takeC}^1, \text{putC}^2, \text{takeS}]
\]

The process in \( \pi \)-calculus notation is \( \pi z. \pi a. \text{Stop} \mid \pi w. \pi a. \text{Stop} \mid x(y). y(u) \), which is should-convergent. We write the variable name of the value-MVar also at the check-MVar commands. Note that after a \( \text{takeC}_x y \), the \( y \) in the rest of the thread is changed accordingly.

The program consists of three threads, where the sequence of the problematic execution is:

\[
\begin{array}{|c|c|c|c|c|c|}
\hline
\text{send} & \text{receive} \\
\hline
[\text{putS}_x z, \text{putC}_x^1, \text{takeC}_x^2, \text{putS}_z a, \text{putC}_z^1, \text{takeC}_z^2, \text{Stop}] & [\text{takeC}_x^1, \text{putC}_x^2, \text{takeS}_x y, \text{takeC}_y^1, \text{putC}_y^2, \text{takeS}_y u] \\
\hline
1 & 2 & 3 & 4 & 5 & 6 & 7 & 8 & 9 & 10
\end{array}
\]

The final program is deadlocked but not successful, hence this translation is not invariant for should-convergence.
Embedding the Pi-Calculus into CH

\[ \psi_{0,T}(P) = \nu \text{stop} \cdot (\text{putMVar stop } \mid \text{stop m } \mid \psi_T(P)) \]

\[ \psi_T(\nu x. P) = \nu x, \text{contx}, \text{checkx}_1, \ldots, \text{checkx}_n \]

\[ (\text{contx m } \mid \text{checkx}_1 m \mid \ldots \mid \text{checkx}_n m \mid x = \text{Channel contx checkx}_1 \ldots \text{checkx}_n \mid \psi_T(P)) \]

\[ \psi_T(\tau y. P) = \epsilon \phi_T(\tau y. P) \]

\[ \psi_T(x(y). P) = \epsilon \phi_T(x(y). P) \]

\[ \psi_T(0) = \epsilon \text{return } () \]

\[ \psi_T(\text{Stop}) = \epsilon \text{takeMVar stop} \]

\[ \psi_T(P | Q) = \psi_T(P) | \psi_T(Q) \]

\[ \psi_T(!P) = \nu f. (\epsilon f \mid \text{do } \{ \text{forkIO } \phi_T(P); f \}) \]

Fig. 10. Induced translations \( \psi_T \) and \( \psi_{0,T} \) for a gstb-translation \( T \) using \( n \) check-MVars

B.2 Correctness of Translations with Global Names

In the correctness proof of translation \( \tau \) and \( \tau_0 \), we defined the top-down translation \( \sigma, \sigma_0 \). For showing correctness of induced translations \( \phi_T \), we follow a similar way, and define induced top-down translations \( \psi_T \) and \( \psi_{0,T} \) in Fig. [10] for a given gstb-translation \( T = (T_{\text{send}}, T_{\text{receive}}) \) using \( n \) check-MVars.

We first show the strong relation between \( \phi_T \) and \( \psi_T \):

**Proposition B.1.** For all \( P \in \Pi_{\text{Stop}} \)

\begin{align*}
&\text{i)} \phi_{0,T}(P) \xrightarrow{sr, \sigma} \psi_{0,T}(P), \quad \text{ii)} \epsilon \phi_T(P) \xrightarrow{sr, \sigma} \psi_T(P), \\
&\text{iii)} \psi_T(P) \sim_c \epsilon \phi_T(P) \text{ and} \quad \text{iv)} \phi_{0,T}(P) \sim_c \psi_{0,T}(P).
\end{align*}

**Proof.** Parts i) and ii) can be proved similar to Lemma [A.3]. Part i) is completely analogous, while for part ii) the translation of the \( \nu \)-operator is different, and thus the construction in the induction for this case is

\[ \epsilon \phi_T(\nu x. P) \xrightarrow{sr,nmvar} \epsilon \phi_T(x) \xrightarrow{sr,beta} \psi_T(P) \]

where \( n \) is the number of check-MVars and where all \( (sr, tmvar) \)-steps are deterministic. Parts iii) and iv) follow from Proposition [B.6] analogously to the proof of Proposition [A.4].

Now, we can prove preservation of may-convergence for any induced translation of an executable gstb-translation, and preservation of may-divergence, provided that must-divergence equivalence already holds:

**Proposition B.2.** Let \( T \) be an executable gstb-translation. Let \( P \) be a closed \( \Pi_{\text{Stop}} \)-process.

- If \( P \downarrow \) then \( \phi_{0,T}(P) \downarrow \).
- If for all closed \( Q \in \Pi_{\text{Stop}}: Q \uparrow \iff \psi_{0,T}(Q) \uparrow \) then the implication \( P \uparrow \implies \phi_{0,T}(P) \uparrow \) holds.
Proof. The proof of the first item is analogous to the proof of Proposition A.9 by an induction on the given converging reduction sequence for \( P \), that constructs a converging sequence for \( \psi_{0,T}(P) \) and then applies Proposition B.1 to derive the sequence for \( \phi_{0,T}(P) \). The second item follows also by an induction on the given reduction sequence for \( P \) to a must-divergent process, where the base case is covered by the precondition.

We sketch the main steps, required to perform the induction:

- For the base case, one can verify that if a closed process \( P \) of the calculus \( \Pi_{\text{Stop}} \) is successful, then \( \psi_{0,T}(P) \xrightarrow{sr,*} Q \) where \( Q \) is successful.

- For the translation of \( \text{dia} \rightarrow \)-steps the following claim hold:
  
  Let \( P \in \Pi_{\text{Stop}} \) be a closed process with \( P \xrightarrow{\text{dia}} P' \). Then there is a sequence \( \psi_{0,T}(P) \xrightarrow{sr,*} \psi_{0,T}(P') \) in the CH-calculus.

  The proof is similar to the proof of Lemma A.5 where one can verify that
  
  - the sr-reduction for the CH-program belonging to \( \text{putS} \) or \( \text{putC} \) consists of
    \[ \xrightarrow{sr,cpcx} \xrightarrow{sr,case} \xrightarrow{sr,pmvar} \beta \]

  - the sr-reduction for the CH-program belonging to \( \text{takeS} \) or \( \text{takeC} \) consists of
    \[ \xrightarrow{sr,cpcx} \xrightarrow{sr,case} \xrightarrow{sr,takeMVar} \beta \]

  and that no garbage (in difference to Lemma A.5) is generated, and that after executing sender and receiver program the requirements on an executable \( \text{gstb} \)-translation imply that the corresponding MVars are again empty.

- For other reduction steps in the \( \Pi_{\text{Stop}} \)-calculus the following claim can be proved analogous to Lemma A.6 with a different translation of \( \nu x \), but this does not change the result:

  Let \( P, P' \in \Pi_{\text{Stop}} \) be closed, such that \( P \xrightarrow{\text{dsc}} P' \). Then \( \psi_{0,T}(P) \xrightarrow{sr,*} \psi_{0,T}(P') \), if rule (replunfold) is used, and \( \psi_{0,T}(P) \equiv \psi_T(P') \), otherwise. In particular, \( \psi_{0,T}(P) \sim c \psi_{0,T}(P') \).

To prove convergence equivalence for \( \phi_{0,T} \) for a given \( \text{gstb} \)-translation it thus suffices to show, that \( \psi_{0,T} \) reflects may-convergence and may-divergence.

### B.3 Correctness of a Translation with the Interprocess Check Restriction and Using Three Global Check-MVars

We sketch that the first translation in Table 2 is correct. The translation is

\[ \xi_1 = (T_{\text{send}}, T_{\text{receive}}) = ([ \text{putS, putC}^1, \text{takeC}^2, \text{putC}^3], [ \text{takeC}^1, \text{putC}^2, \text{takeC}^3, \text{takeS}]) \]

We conjecture that all four translations are correct using similar arguments.

We leave open whether the translations 5,6 of Table 2 are correct. In the proof we mention only the reduction steps that correspond to actions.

**Theorem B.3.** For all closed \( \pi \)-processes \( P \) and the translation \( \phi_{0,\xi_1} \), the equivalences \( P_\uparrow \iff \phi_{0,\xi_1}(P)_\downarrow \) and \( P_\downarrow \iff \phi_{0,\xi_1}(P)_\downarrow \) hold.

**Proof.** 1. If \( P_\uparrow \), then \( \phi_{0,\xi_1}(P)_\downarrow \) follows from Proposition B.2
2. Let us assume that \( \psi_{0, \tau_1}(P) \downarrow \), and let \( \psi_{0, \tau_1}(P) \xrightarrow{\sigma} s_0 \) be a successful reduction sequence. The task is now to find a rearrangement of the reduction sequence which remains successful, and that can be translated back into a \( \Pi_{St} \)-calculus reduction sequence. We only consider the MVar accesses ignoring the other reduction steps and try to make an appropriate rearrangement. We can assume that there is an access to an MVar, say \( x \), in the reduction sequence. Due to the initialization, the first one is a putMVar on \( x \).

If the sequence \( \text{putC}_x^1, \text{takeC}_x^2, \text{putC}_x^3 \) in the thread of \( \text{putS}_x \) is not finished, then this thread does not contribute to the success. Also the \( T_{\text{receive}} \)-sequence cannot be finished. Hence we can remove all the reductions from both MVar-access-sequences without changing the success. This makes the reduction sequence smaller and we can use induction. If the \( T_{\text{send}} \)-sequence is complete, but the \( T_{\text{receive}} \)-sequence is missing, then it is impossible that another \( T_{\text{receive}} \)-sequence starts (see Proposition 5.8). Hence we can add the \( \text{takeC}_x^3, \text{takeS}_x \)-sequence or only the action \( \text{takeS}_x \) without changing success and thus the two sequences are completely done, and retranslatable. We can move the reductions as far to the left as possible, close to the start of \( T_{\text{send}} \) with \( \text{putS}_x \). Then we can apply induction on the remaining part of the reduction sequence.

Thus we have \( \psi_{0, \tau_1}(P) \downarrow \implies P \downarrow \). By Proposition B.1 this also shows \( \phi_{0, \tau_1}(P) \downarrow \implies P \downarrow \).

3. Now we can assume that \( P \downarrow \iff \psi_{0, \tau_1}(P) \downarrow \), and also \( P \uparrow \iff \psi_{0, \tau_1}(P) \uparrow \).

4. The implication \( P \uparrow \implies \phi_{0, \tau_1}(P) \uparrow \) now follows from Proposition B.2.

5. The implication \( \phi_{0, \tau_1}(P) \uparrow \implies P \uparrow \) is a bit more complex and requires again rearranging the reduction sequence. We show \( \psi_{0, \tau_1}(P) \uparrow \implies P \uparrow \) and apply Proposition B.3 to derive the result for \( \phi_{0, \tau_1}(P) \uparrow \).

However, we now have to argue more in the case of reduction subsequences that do not correspond to a translated interact. Therefore we have to show that the removal of certain reduction steps and/or the addition of reduction steps is invariant w.r.t. must-divergent processes.

Due to Proposition [5,8] it is sufficient to consider a reduction sequence that exhibits may-divergence, and look for the leftmost single MVar accesses:

Let \( \psi_{0, \tau_1}(P) \xrightarrow{\sigma} s_\infty \) be a reduction sequence where \( s_\infty \) is must-divergent.

- Let there be a leftmost \( \text{putS}_x \) for channel \( x \) and also perhaps a subsequent \( \text{putC}_x^3 \), but no more actions for \( x \) in the sequence. Let \( s_2 \) be the process after performing \( \text{putS}_x \) (or \( \text{putC}_x^3 \), if present), and \( s_1 \) be the process before performing \( \text{putS}_x \). Then we distinguish two cases: i) If there is no other thread which can perform \( \text{takeC}_1 \), then we can remove the operations, constructing a \( s'_\infty \) which still must be must-divergent: Any other send-sequence on \( x \) starting with \( \text{putS}_x \) performed by another thread also cannot be completed, since there is no receiver that can perform \( \text{takeC}_1 \). ii) There is another thread which can perform \( \text{takeC}_1 \) (but does not do it). Then this thread is blocked and still present in \( s_\infty \). Thus we can add the step after \( s_\infty \) and derive an \( s'_\infty \) and then we shift the reductions that perform the \( \text{takeC}_1 \) to the left.
– The next case is the execution of \( \text{putS}_x \) and \( \text{putC}_1^x \) in thread 1, and \( \text{takeC}_1^2x \) in thread 2, but no further actions in the reduction sequence. Now it is not possible to argue using removal of the steps, since there may be a may-convergent reduction sequence using another thread for \( \text{takeC}_1^1x \). Now Proposition 5.8 can be used. If we look for a possible reduction sequence from \( s_{\infty} \), then there is one reduction sequence (almost \( sr \)-reduction sequence if we ignore the main thread) that executes the complete interact for \( x \). Since \( s_{\infty} \) is must-divergent, we can add reduction steps and the result will be must-divergent. Hence we can prolong the reduction sequence and also by shifting the missing reduction to the left obtain a sequence where the translated interact is complete.

– the final case is that we have the 8 actions for channel \( x \) in two threads. These can be shifted to the left, since there is no disturbance with other threads, and we obtain a complete interact that is retranslatable.

B.4 Correctness of a Translation without the Interprocess Check Restriction and Two Global Check-MVars

We sketch that the first translation in Table 3 is correct. The translation is

\[
\mathcal{T}_2 = (T_{\text{send}}, T_{\text{receive}}) = ([\text{putC}_1, \text{putS}, \text{takeC}_2, \text{takeC}_1], [\text{takeS}, \text{putC}_2])
\]

We conjecture that the second translation in Table 3 can be proved correct using similar arguments. We mention only the reductions that correspond to actions.

**Theorem B.4.** For all \( \pi \)-processes \( P \) and the translation \( \phi_0, \mathcal{T}_2 \), the equivalences

\[
P \downarrow \iff \phi_0, \mathcal{T}_2(P) \downarrow \quad \text{and} \quad P \uparrow \iff \phi_0, \mathcal{T}_2(P) \uparrow
\]

hold.

**Proof.** We can reason similar to the proof of Theorem B.3. Concentrating on the core of the proof, it suffices to show that \( \psi_{\mathcal{T}_2}(P) \downarrow \implies P \downarrow \) and \( \psi_{\mathcal{T}_2}(P) \uparrow \implies P \uparrow \), where for the second implication we can assume that \( Q \uparrow \iff \psi_{\mathcal{T}_2}(Q) \uparrow \) holds for all closed \( II_{\text{Stop}} \)-processes \( Q \).

1. Let us assume that \( \psi_{\mathcal{T}_2}(P) \), and let \( \psi_{\mathcal{T}_2}(P) \overset{*}{\rightarrow} s_0 \) be a successful reduction sequence. The task is now to find a rearrangement of the reduction sequence which remains successful, and that can be translated back into a \( II_{\text{Stop}} \)-calculus reduction sequence. We only consider the MVar accesses ignoring the other reduction steps and try to make an appropriate rearrangement. We can assume that there is an access to an MVar, say \( x \), in the reduction sequence. Due to the initialization, the first one is a \( \text{putC}_1^x \). If the sequence \( \text{putS}_x, \text{takeC}_2^x, \text{takeC}_1^x \) in the thread of \( \text{putC}_1^x \) is not finished, then this thread does not contribute to the success. Also the \( T_{\text{receive}} \) sequence cannot be finished. Hence we can remove all the reductions from both MVar-access-sequences without changing the success. This makes the reduction sequence smaller and we can use induction. If the \( T_{\text{send}} \)-sequence is complete, but the \( \text{takeS}_x \) of the \( T_{\text{receive}} \)-sequence is missing, then it is impossible that another \( T_{\text{receive}} \)-sequence starts (see Proposition 5.10) Hence
we can add the \texttt{takeS}, \texttt{putC}^2\texttt{x}-sequence without changing success and thus the two sequences are completely done, and retranslatable. We can move the reductions as far to the left as possible, close to the start of $T_{send}$ with \texttt{putC}^1\texttt{x}.

Then we apply induction on the remaining part of the reduction sequence.

2. The implication $\psi_{T_2}(P)^\uparrow \implies P^\uparrow$ requires again rearranging the reduction sequence. However, we now have to argue more in the case of reduction subsequences that do not correspond to a translated interact. Therefore we have to show that the removal of certain reduction steps and/or the addition of reduction steps is invariant w.r.t. must-divergent processes. Due to Proposition 5.10 it is sufficient to consider a reduction sequence that exhibits may-divergence, and look for the leftmost single MVar accesses: Let $\psi_{T_2}(P) \xrightarrow{sr} s_\infty$ be a reduction sequence where $s_\infty$ is must-divergent.

- Let there be a leftmost \texttt{putC}^1\texttt{x} for channel $x$, and also perhaps a subsequent \texttt{putS}$_x$, but no more actions for $x$ in the sequence. If there is no thread that could perform \texttt{takeS}$_x$, then we can remove the \texttt{putC}^1\texttt{x} and \texttt{putS}$_x$-executions and the adapted process $s'_\infty$ remains to be must-divergent, since any other sender that performs \texttt{putC}^1\texttt{x} also cannot be completed. If there is a thread that can perform \texttt{takeS}$_x$, then this is still possible in $s_\infty$. Thus we add the \texttt{takeS}$_x$ (and the \texttt{putS}$_x$ if needed) for $s_\infty$ and then shift the corresponding reductions to the left. Then we can use another case of this proof.

- The next case is the execution of \texttt{putC}^1\texttt{x} and \texttt{putS}$_x$ in thread 1, and \texttt{takeS}$_x$ in thread 2, but no further actions in the reduction sequence. Now Proposition 5.10 implies that if we look for a possible reduction sequence from $s_\infty$, then there is one reduction sequence (almost sr-reduction sequence if we ignore the main thread) that executes the complete interact for $x$. Since $s_\infty$ is must-divergent, we can add reduction steps and the result will be must-divergent. Hence we can prolong the reduction sequence and also by shifting the missing reduction to the left obtain a sequence where the translated interact is complete.

- The next case is the execution of \texttt{putC}^1\texttt{x} and \texttt{putS}$_x$ (and perhaps \texttt{takeC}^2\texttt{x}) in thread 1, and \texttt{takeS}$_x$, \texttt{putC}^2\texttt{x} in thread 2, but no further actions in the reduction sequence. Again Proposition 5.10 implies that if we look for a possible reduction sequence from $s_\infty$, then there is one reduction sequence (almost sr-reduction sequence if we ignore the main thread) that executes the complete interact for $x$. Since $s_\infty$ is must-divergent, we can add reduction steps and the result will be must-divergent. Hence we can prolong the reduction sequence and also by shifting the missing reduction to the left obtain a sequence where the translated interact is complete.

- the final case is that we have the 6 actions for channel $x$ in two threads. These can be shifted to the left, since there is no disturbance with other threads, and we obtain a complete interact that is retranslatable.
Proposition 3.5. Let $P_1, P_2$ be well-formed and $P_1 \equiv P_2$. Then $P_1 \sim_c P_2$.

Proof. Let $P_1, P_2$ be well-formed, $P_1 \equiv P_2$, and $D \in PCtx_{CH}$. We have to show i) $D[P_1] \downarrow \Rightarrow D[P_2] \downarrow$; ii) $D[P_2] \downarrow \Rightarrow D[P_1] \downarrow$; iii) $D[P_1] \downarrow \Rightarrow D[P_2] \downarrow$; and iv) $D[P_2] \downarrow \Rightarrow D[P_1] \downarrow$. The process $D[P_1]$ is successful iff $D[P_2]$ is successful, since $\equiv$ cannot remove or introduce a main-thread. For any process $P_3$, we have:
Embedding the Pi-Calculus into CH

Let $P_1 \xrightarrow{sr} P_3$ iff $P_2 \xrightarrow{sr} P_3$, since $\equiv$ is a congruence and since $\rightarrow$ is closed w.r.t. $\equiv$. Using both facts, we can use induction on a given reduction sequence starting with $\mathbb{D}[P_i]$ and ending in a successful process, to show that $\mathbb{D}[P_j]$ reduces to a successful process, where $(i, j) \in \{(1, 2), (2, 1)\}$. This shows parts i) and ii).

For the remaining parts, we show $\mathbb{D}[P_1] \uparrow \iff \mathbb{D}[P_2] \uparrow$. We use induction on a given reduction sequence starting with $\mathbb{D}[P_i]$ and ending in must-divergent process to show that $\mathbb{D}[P_j] \uparrow$ where $(i, j) \in \{(1, 2), (2, 1)\}$. The base case holds, since parts i) and ii) imply $\mathbb{D}[P_1] \uparrow \iff \mathbb{D}[P_2] \uparrow$ and the induction step again uses the fact, that each reduction step applied to $\mathbb{D}[P_i]$ can also be applied to $\mathbb{D}[P_j]$ for $(i, j) \in \{(1, 2), (2, 1)\}$.

C.3 Embedding CH in CHF

We briefly recall the program calculus CHF [22,23] and then provide an embedding of CH into CHF. After showing that the embedding is adequate w.r.t. contextual semantics we transfer results on the correctness of program transformations from CHF to CH. Since the calculus CHF is quite similar to the calculus CH, we often only comment on the differences between both calculi.

The syntax of processes, expressions, monadic expressions and types is given in Fig. 12. The laws of structural congruence and the syntax of contexts are given in Figs. 13 and 14. The calculus CHF uses the same types as the calculus CH. Expressions in CHF are like expressions in CH, with the only difference that the monadic operator forkIO is replaced by the monadic operator future. While forkIO is of type IO () → IO () the operator future is of type IO t → IO t for any type t. On the process-level CH threads $\equiv e$ are replaced by futures $x \equiv e$ where x is an introduced variable, called the name of the future. The operational behavior is, that expression e (of type IO t) is evaluated concurrently and thereafter the result (of type t) is bound to variable x, which makes the value of the concurrent computation accessible. In the standard reduction of the calculus CHF (see Fig. 15) this is implemented by the following changes w.r.t. the standard reduction in CH: the rule (fork) creates the new future (instead of a new thread) and the calling thread receives the name of the future as result. If a concurrent thread finishes its computation, then the result is shared as a global binding and the thread is removed by the (new) rule (unIO).

The main thread in CHF is a future and it is written as $x \equiv e$. Definitions of functional values, monadic values, and values are as in CH adapted to the syntax changes. A well-formed process $P$ is successful, if $P \equiv \nu x_1.\ldots.\nu x_n.(x \equiv e \equiv \text{return } e \equiv P')$. The definitions of may-convergence and should-convergence in CHF are the same as in CH [see Definition 3.2], but adapted to the syntax changes and using the standard reduction defined in Fig. 15. Contextual approximation $\leq_c$ and contextual equivalence $\sim_c$ on CHF-processes and expressions are defined analogous to Definition 3.3. Well-formed and structural congruent CHF-processes $P_1 \equiv P_2$ are also contextual equivalent (i.e. $P_1 \sim_c P_2$, see [22]).
\[ P \in \text{Proc}_{\text{CHF}} \triangleq (P_1 \mid P_2) \mid x \leftarrow e \mid \nu x.P \mid x \cdot m \cdot x = e \]

\[ e \in \text{Expr}_{\text{CHF}} \triangleq x \mid m \mid \lambda x.e \mid (e_1 \cdot e_2) \mid \text{seq} \ e_1 \cdot e_2 \mid e \ e_1 \ldots e_n \]

\[ \text{letrec } x_1 = e_1, \ldots, x_n = e_n \text{ in } e \]

\[ \text{case } e \text{ of } (c_{T_1} \cdot x \ldots x_{ar(c)} \rightarrow e_1) \ldots (c_{T_n} \cdot x_1 \ldots x_{ar(c_T)} \rightarrow e_T) \]

\[ m \in \text{MExml}_{\text{CHF}} \triangleq \text{return } e \mid e \Rightarrow e' \mid \text{future } e \mid \text{takeMVar } e \mid \text{newMVar } e \mid \text{putMVar } e \ e' \]

\[ t \in \text{Typ}_{\text{CHF}} \triangleq 10 \ t \mid (T \ t_1 \ldots t_n) \mid \text{MVar } t \mid t_1 \rightarrow t_2 \]

Fig. 12. Syntax of expressions, processes, and types of CHF

\[ P_1 \mid P_2 \equiv P_3 \mid P_1 \]

\[ (P_1 \mid P_2) \mid P_3 \equiv P_4 \mid (P_2 \mid P_3) \]

\[ (\nu x.\ P_1) \mid P_2 \equiv (\nu x.\ P_1) \mid P_2 \text{ if } x \notin FV(P_2) \]

\[ \nu x_1 \nu x_2.\ P \equiv \nu x_1.\ (\nu x_2.\ P) \]

\[ P_1 \equiv P_2 \text{ if } P_1 =_\alpha P_2 \]

Fig. 13. Structural congruence of CHF

\[ \mathbb{D} \in \text{PCxt}_{\text{CHF}} \triangleq [] \mid \mathbb{D} \mid P \mid P \mid P \mid \nu x.\ \mathbb{D} \]

\[ M \in \text{MCxt}_{\text{CHF}} \triangleq [] \mid M \Rightarrow e \]

\[ \mathbb{F} \in \text{FCxt}_{\text{CHF}} \triangleq E \mid (\text{takeMVar } E) \mid (\text{putMVar } E) \]

\[ E \in \text{ECxt}_{\text{CHF}} \triangleq [] \mid (E) \mid (\text{seq } E) \mid (\text{case } E \text{ of } alts) \]

Fig. 14. PCxt_{\text{CHF}}, MCxt_{\text{CHF}}, FCxt_{\text{CHF}}, ECxt_{\text{CHF}}-contexts

Monadic Computations

- (lunit) \[ y \equiv M[\text{return } e_1 \Rightarrow e_2] \xrightarrow{s} y \equiv M[e_2 e_1] \]
- (tmvar) \[ y \equiv M[\text{takeMVar } x] \mid x \cdot m \equiv y \leftarrow M[\text{return } e] \mid x \cdot m \equiv \]
- (pmvar) \[ y \equiv M[\text{putMVar } x] \mid x \cdot m \equiv y \leftarrow M[\text{return } ()] \mid x \cdot m \equiv \]
- (nmvar) \[ y \equiv M[\text{newMVar } e] \xrightarrow{s} \nu z.\ (y \leftarrow M[\text{return } x] \mid z \cdot e) \]
- (fork) \[ y \equiv M[\text{future } e] \xrightarrow{s} \nu z.\ (y \leftarrow M[\text{return } z] \mid z \cdot e) \text{ where } z \text{ is fresh} \]
- (unIO) \[ y \equiv \text{return } e \xrightarrow{s} y \equiv e \text{ if the thread is not the main-thread} \]

Functional Evaluation

- (ccpe) \[ y \equiv M[\text{bind } x = e] \xrightarrow{s} y \equiv M[\text{bind } e] \mid x = e \]
- (mkbinds) \[ y \equiv M[\text{letrec } x_1 = e_1, \ldots, x_n = e_n \text{ in } e] \]
- (beta) \[ y \equiv M[\text{case } (\lambda x.\ e_1)] \xrightarrow{s} y \equiv M[\text{bind } e_1] \mid x = e_1 \ldots x_n = e_n] \]
- (case) \[ y \equiv M[M[\text{case } (\lambda x.\ e_1) \ldots (\lambda x.\ e_n) \text{ of }]} \ldots (\text{of } c \ldots c_n) \mid x_1 \ldots x_n] \]
- (seq) \[ y \equiv M[M[\text{seq } e] \xrightarrow{s} y \equiv M[\text{seq } e] \text{ where } v \text{ is a functional value} \]

Closure: If \[ P_1 \equiv \mathbb{D}[P'_1], P_2 \equiv \mathbb{D}[P'_2], \] and \[ P'_1 \xrightarrow{s} P'_2 \] then \[ P_1 \xrightarrow{s} P_2. \]

Capture avoidance: We assume capture avoiding reduction for all reduction rules.

Fig. 15. Standard reduction rules of CHF (call-by-name-version)
(gc) \( \nu x_1, \ldots, x_n. (P \mid \text{Comp}(x_1) \mid \ldots \mid \text{Comp}(x_n)) \rightarrow P \),
if \( \forall i \in \{1, \ldots, n\} : \text{Comp}(x_i) \) is a binding \( x_i = e_i \), an MVar \( x_i \), \( m \), or \( x_i = - \),
and \( x_1, \ldots, x_n \not\in \text{FV}(P) \)

(dtmvar) \( \nu x. \mathcal{D}[y \leftarrow M[\text{takeMVar } x] \mid x \cdot m \cdot e] \rightarrow \nu x. \mathcal{D}[y \leftarrow M[\text{return } e] \mid x \cdot m \cdot -] \),
if for \( \forall \mathcal{D}' \in \text{PCTxt}_{\text{CHF}} \) the first execution of \( \text{takeMVar } x \) is in thread \( y \) for
all sr-reductions starting with \( \mathcal{D}'[\nu x. (\mathcal{D}[y \leftarrow M[\text{takeMVar } x] \mid x \cdot m \cdot e]) \]

(dpmpvar) \( \nu x. \mathcal{D}[y \leftarrow M[\text{putMVar } x \cdot e] \mid x \cdot m \cdot -] \rightarrow \nu x. \mathcal{D}[y \leftarrow M[\text{return } ()] \mid x \cdot m \cdot e] \),
if \( \forall \mathcal{D}' \in \text{PCTxt}_{\text{CHF}} \) and any \( e' \) the first execution of \( \text{putMVar } x \cdot e' \) for all sr-
reductions starting with \( \mathcal{D}'[\nu x. (\mathcal{D}[y \leftarrow M[\text{putMVar } x \cdot e] \mid x \cdot m \cdot -]) \]

Fig. 16. Transformations (gc), (dtmvar), and (dpmpvar)

Program transformations \( \eta \) in \( \text{CHF} \) are binary relations between \( \text{CHF} \)-processes,
where \( \eta \) is correct iff \( \eta \subseteq \sim_e \). We define garbage collection (gc) and deterministic
variants of (tmvar) and (pmvar) for the calculus \( \text{CHF} \) in Fig. [16]

**Proposition C.1** ([22]). The transformations (lunit), (nmvar), (fork), (unIO),
(cpce), (mkbinds), (beta), (case), (seq) (gc), (dtmvar), and (dpmpvar) are correct
program transformations, whereas (pmvar) and (tmvar) in general are not correct
as program transformations.

We define an embedding \( \iota \) which embeds \( \text{CH} \) into \( \text{CHF} \):

**Definition C.2.** We define the translation \( \iota : \text{CH} \rightarrow \text{CHF} \). For most of the
cases \( \iota \) is applied homomorphically over the term structure, without changing the
syntax, i.e. for variables \( r \) and nullary constructors \( r : \iota(r) := r \) and for any syn-
tactic operator \( f \) and arguments \( r_1, \ldots, r_n \) : \( \iota(f \cdot r_1 \cdot \ldots \cdot r_n) := f(\iota(r_1) \cdot \ldots \cdot \iota(r_n)) \).
We now explicitly list the exceptional cases: For expressions \( \text{forkIO } e \) we define:

\[
\iota(\text{forkIO } e) := (\text{future } \iota(e)) \gg \text{return } ()
\]

The exceptional cases on the process level are

\[
\iota(\leftarrow e) := \nu x. x \leftarrow \iota(e) \quad \text{and} \quad \iota(\leftarrow e) \leftarrow \nu x. x \leftarrow \iota(e)
\]

where for both cases, we assume that \( x \) is chosen (fresh) such that \( x \) does not occur in \( \iota(e) \). On types, the translation \( \iota \) is the identity.

**Proposition C.3.** Assume that the countable infinite set of variables in \( \text{CH} \) is
\( \text{Var}_{\text{CH}} \) and in \( \text{CHF} \) is \( \text{Var}_{\text{CHF}} = \text{Var}_{\text{CH}} \cup \text{Var}' \) where \( \text{Var}_{\text{CH}} \cap \text{Var}' = \emptyset \), \( \text{Var}' \) is
countably infinite and translation \( \iota \) always uses a variable of \( \text{Var}' \) for translating
\( \leftarrow e \) or \( \leftarrow e \). Then the translation \( \iota \) is compositional, i.e. \( \iota(C[e]) = \iota(C)[\iota(e)] \).

For analyzing the correspondence between the sr-reductions in \( \text{CH} \) and \( \text{CHF} \)
w.r.t. \( \iota \), an observation is that for all \( \mathcal{D} \in \text{PCTxt}_{\text{CH}} : \iota(\mathcal{D}) \in \text{PCTxt}_{\text{CHF}} \); for all \( M \in \text{MCtx}_{\text{CH}} : \iota(M) \in \text{MCtx}_{\text{CHF}} \); for all \( E \in \text{ETxt}_{\text{CH}} : \iota(E) \in \text{ETxt}_{\text{CHF}} \); and for all \( F \in \text{FTxt}_{\text{CH}} : \iota(F) \in \text{FTxt}_{\text{CHF}} \). We analyze the sr-reduction w.r.t. \( \iota \).
Lemma C.4. Let $P, P'$ be CH-processes with $P \xrightarrow{sr} P'$. Then either $\iota(P) \xrightarrow{sr,0} \iota(P')$ or $\iota(P) \xrightarrow{sr,3} \iota(P')$ where in the latter case $\iota(P) \sim_c \iota(P')$.

Proof. We consider the different reduction rules. For rules (lunit), (tmvar), (pmvar), and for all functional evaluations, we have: if $P \xrightarrow{sr} P$ then $\iota(P) \xrightarrow{sr} \iota(P')$. For (fork), let $P = \mathbb{D}[\leftarrow \mathbb{M}[\text{fork}IO e]] \xrightarrow{sr} \mathbb{D}[\leftarrow \mathbb{M}[\text{return }] | l \leftarrow e] = P'$.

Then we can reduce $\iota(P)$ as follows:

$$
\iota(P) = \iota(\mathbb{D})[\nu x.x \leftarrow \text{future } \iota(e) \gg \lambda_{\text{return }} ()] \\
\xrightarrow{sr,\text{fork}} \iota(\mathbb{D})[\nu x.\nu z.x \leftarrow \text{return } z \gg \lambda_{\text{return }} () | z \leftarrow \iota(e)] \\
\xrightarrow{sr,\text{unit}} \iota(\mathbb{D})[\nu x.\nu z.x \leftarrow (\lambda_{\text{return }} () | z \leftarrow \iota(e))] \\
\xrightarrow{sr,\text{beta}} \iota(\mathbb{D})[\nu x.\nu z.x \leftarrow \text{return } () | z \leftarrow \iota(e)] = \iota(P')
$$

Theorem C.1 implies $\iota(P) = \iota(P')$.

Lemma C.5. $P \in \text{Proc}_{\text{CH}}$ is successful iff $\iota(P)$ is successful.

Proof. If $P \in \text{Proc}_{\text{CH}}$ is successful, then $P \equiv \mathbb{D}[\leftarrow \mathbb{M}[\text{return }] |]$ and $\iota(P) = \iota(\mathbb{D})[\nu x.x \leftarrow \text{main }] \equiv \text{return } ()$. If $\iota(P)$ is successful, then $\iota(P) = \mathbb{D}[x \equiv \text{return } ()]$ and thus $P \equiv \mathbb{D}'[\leftarrow \text{return } ()]$ for some $\mathbb{D}' \in \text{Ptxt}_{\text{CH}}$.

Proposition C.6. Let $P \in \text{Proc}_{\text{CH}}$ such that $P \downarrow$. Then $\iota(P) \downarrow$.

Proof. By induction on $P \xrightarrow{sr} P'$ where $P'$ is successful. The base case holds by Lemma C.5. For the induction step, let $P \xrightarrow{sr} P'' \xrightarrow{sr,a} P'$. The induction hypothesis shows $\iota(P'') \downarrow$ and Lemma C.4 shows $\iota(P) \xrightarrow{sr,a} \iota(P'')$. Thus $\iota(P) \downarrow$.

Checking all cases of the standard reduction in CHF together with inspecting the images of the translation $\iota$ shows:

Lemma C.7. Let $P \in \text{Proc}_{\text{CH}}$, $\iota(P) \xrightarrow{sr,a} P'$, and $a \notin \{\text{fork, unIO}\}$. Then there exist $P''$ with $P \xrightarrow{sr} P''$ such that $\iota(P'') = P'$, i.e. the diagram shown on the right side holds.

Lemma C.8. Let $P \in \text{Proc}_{\text{CH}}$ such that $\iota(P) \xrightarrow{sr,\text{fork}} P'$. Then there exists a CH-process $P''$ where $P \xrightarrow{sr,\text{fork}} P'' \xrightarrow{sr,a} P'$ and $P' \xrightarrow{sr,a} \iota(P'')$. The equivalence $\iota(P'') \sim_c P'$ holds. The diagram on the right hand side shows the given and existentially quantified reductions and translations steps.

Proof. The (sr,fork)-step in CHF evaluates a future-operation that must stem from translating a forkIO. Since $\iota(\text{fork}IO e) = \text{future } \iota(e) \gg \text{return } ()$, the reduction $\iota(P) \xrightarrow{sr,\text{fork}} P'$ can be extended by two reduction steps $P' \xrightarrow{sr,\text{unit}} P''$ such that $P \xrightarrow{sr,\text{fork}} P''$ and $\iota(P'') = P''$. Since (sr,beta) and (sr,fork) are correct in CHF (Proposition C.4), we have $P'' \sim_c P'$. □
The two reductions leading from \( P' \) to \( \iota(P'') \) in the previous lemma, are correct program transformations, and they are deterministic for the corresponding thread (i.e., for this thread, no other reduction is possible, even if other components or threads change (for instance, the content of MVars). Thus, we can safely add all the missing reduction steps corresponding to Lemma C.8 to a given sequence \( \iota(P) \xrightarrow{sr,unIO} P_0 \) resulting in a sequence \( \iota(P) \xrightarrow{sr,*} P_1 \) such that i) \( P_0 \sim_c P_1 \) and ii) \( P_1 \) is successful if \( P_0 \) is successful.

Lemmas C.7 and C.8 do not cover the case where \( \iota(P) \xrightarrow{sr,unIO} P' \) for some \( P \in \text{Proc}_{CH} \). In this case \( P = D[\epsilon \Leftarrow \text{return } ()] \) must hold, and \( \iota(P) = \iota(D)[\nu x.x \Leftarrow \text{return } ()] \xrightarrow{sr,unIO} \iota(D)[\nu x.x = ()] \). Since future \( x \) cannot be used in the context \( \iota(D) \), it is irrelevant, if the \( (sr,unIO) \)-reduction is performed or is not performed. I.e., given a sequence \( \iota(P) \xrightarrow{sr,n} P_1 \), we can modify the sequence, by removing all \( (sr,unIO) \)-reductions, resulting in a sequence \( \iota(P) \xrightarrow{sr,*} P_2 \) such that i) \( P_1 \sim_c P_2 \) and ii) \( P_1 \) is successful if and only if \( P_2 \) is successful.

**Definition C.9.** Let \( \text{Red} = \iota(P) \xrightarrow{sr,*} P' \) be a CHF-standard reduction sequence for \( P \in \text{Proc}_{CH} \). We say that \( \text{Red} \) is \( \iota \)-normalized if the two described transformations on reduction sequences (adding reduction steps corresponding to Lemma C.8 and removing \( (sr,unIO) \)-reductions) are not applicable to \( \text{Red} \).

As argued above, the following lemma holds:

**Lemma C.10.** For every sequence \( \iota(P) \xrightarrow{sr,*} P' \) with \( P \in \text{Proc}_{CH} \), there exists a \( \iota \)-normalized reduction sequence \( \iota(P) \xrightarrow{sr,n} P'' \) such that i) \( P' \sim_c P'' \) and ii) \( P'' \) is successful if \( P' \) is successful.

**Proposition C.11.** Let \( P \in \text{Proc}_{CH} \) and \( \iota(P) \downarrow \). Then \( P \downarrow \).

**Proof.** Let \( \iota(P) \xrightarrow{sr,*} P_0 \) where \( P_0 \) is successful. By Lemma C.10 there exists a \( \iota \)-normalized reduction sequence \( \iota(P) \xrightarrow{sr,n} P_1 \) where \( P_1 \) is successful and \( n \geq 0 \). By induction on \( n \) we show \( P \downarrow \). Lemma C.5 covers the case \( n = 0 \). If \( n > 0 \) then \( \iota(P) \xrightarrow{sr,n} P_1 \). Since the given sequence is \( \iota \)-normalized, we can either apply the commutation diagram from Lemma C.7 or the diagram from Lemma C.8 to a prefix of \( \iota(P) \xrightarrow{sr,n} P_1 \) such that \( P \xrightarrow{sr} P'' \) and \( \iota(P'') \xrightarrow{sr,n'} P_0 \) where \( n' < n \) and \( \iota(P') \xrightarrow{sr,n'} P_1 \) is \( \iota \)-normalized and thus we can apply the induction hypothesis which shows \( P'' \downarrow \) and thus \( P \downarrow \). 

By Propositions C.6 and C.11 we also have:

**Corollary C.12.** Let \( P \in \text{Proc}_{CH} \). Then \( \iota(P) \uparrow \iff P \uparrow \).

**Proposition C.13.** Let \( P \in \text{Proc}_{CH} \) such that \( P \uparrow \). Then \( \iota(P) \uparrow \).

**Proof.** By induction on length \( n \) of a reduction \( P \xrightarrow{sr,n} P' \) where \( P' \uparrow \). Corollary C.12 covers the base case. For the induction step, let \( P \xrightarrow{sr} P'' \xrightarrow{sr,n} P' \). The induction hypothesis shows \( \iota(P'') \uparrow \). Lemma C.4 shows that \( \iota(P) \xrightarrow{sr,*} \iota(P'') \), and thus \( \iota(P) \uparrow \).
Proposition C.14. Let \( P \in \text{Proc}_{CH} \) and \( \iota(P) \uparrow \). Then \( P \uparrow \).

Proof. Let \( \iota(P) \xrightarrow{sr,n} P_0 \) where \( P_0 \uparrow \). By Lemma C.10 there exists a \( \iota \)-normalized reduction sequence \( \iota(P) \xrightarrow{sr,n} P_1 \) where \( P_1 \uparrow \) and \( n \geq 0 \). We use induction on \( n \) to show \( P \uparrow \).

If \( n = 0 \), then Corollary C.12 shows that \( P \uparrow \).

If \( n > 0 \), then \( \iota(P) \xrightarrow{sr,n} P_1 \). Since the sequence is \( \iota \)-normalized, we can either apply the composition diagram from Lemma C.7 or the diagram from Lemma C.8 to a prefix of \( \iota(P) \xrightarrow{sr,n} P_1 \) such that \( P \xrightarrow{sr} P'' \) and \( \iota(P'') \xrightarrow{sr,n'} P_0 \) where \( n' < n \) and also \( \iota(P'') \xrightarrow{sr,n'} P_1 \) is \( \iota \)-normalized and thus we can apply the induction hypothesis which shows \( P'' \uparrow \) and thus also \( P \uparrow \).

\( \square \)

Propositions C.6, C.11, C.13, and C.14 show:

Theorem C.15. The translation \( \iota \) is convergence equivalent.

Theorem C.16. The translation \( \iota \) is adequate, i.e. for all processes \( P_1, P_2 \in \text{Proc}_{CH} \): \( \iota(P_1) \sim_c \iota(P_2) \implies P_1 \sim_c P_2 \) and for all expressions \( e_1, e_2 \in \text{Expr}_{CH} \): \( \iota(e_1) \sim_c \iota(e_2) \implies e_1 \sim_c e_2 \).

Proof. This holds, since \( \iota \) is compositional and convergence equivalent. We only show the part for processes. Let \( \xi \in \{\downarrow, \uparrow\} \), \( (i,j) \in \{(1,2), (2,1)\} \). Let \( \iota(P_i) \sim_c \iota(P_j) \) and \( D \in \text{PCtxt}_{CH} \) such that \( D[P_i]\xi \). Then \( \iota(D)[P_i] = \iota(D)[\iota(P_i)] \) (since \( \iota \) is compositional) and thus \( \iota(D)[\iota(P_i)]\xi \) (since \( \iota \) is convergence equivalent).

Since \( \iota(P_i) \sim_c \iota(P_j) \), this implies \( \iota(D)[\iota(P_j)]\xi \) which implies \( \iota(D)[P_j]\xi \) (since \( \iota \) is compositional) and thus \( D[P_j]\xi \) (since \( \iota \) is convergence equivalent).

\( \square \)

Proposition 3.6. The transformations \( \text{(lunit), (nmvar), (fork), (cpce), (mkbinds), (beta), (case), (seq), (yc), (dtmvar), and (dpmvar) are correct in CH.} \)

Proof. For all mentioned transformations \( \eta \), we have: if \( P \xrightarrow{\eta} P' \) then \( \iota(P) \sim_c \iota(P') \) by Proposition C.1 which implies \( P \sim_c P \) by Theorem C.16.

\( \square \)

D Barbed Convergence Testing and Equivalence

In [24], the following claim is proved, where \( \sigma \) are name-to-name substitutions:

Theorem D.1 ([24]). For all processes \( P, Q \in I_{\text{Stop}} \):

- If for all \( \sigma, R: \sigma(P) \uparrow R \leq_1 \sigma(Q) \uparrow R \), then \( P \leq_c Q \).
- If for all \( \sigma, R: \sigma(P) \uparrow R \leq_\uparrow \sigma(Q) \downarrow R \leq_\uparrow \sigma(Q) \uparrow R \), then \( P \leq_c Q \).

Example D.2. We show the equivalence \( x(y).\text{Stop} \uparrow x(z).\text{Stop} \sim_c \text{Stop} \). Clearly for all \( \sigma \) and all processes \( R \) we have \( \sigma(\text{Stop}) \uparrow R \) is successful and thus \( \sigma(\text{Stop}) \downarrow R \).

We have that \( \sigma(x(y).\text{Stop} \uparrow x(z).\text{Stop}) \downarrow R \), since the process reduces in one step to \( \text{Stop} \uparrow R \). We observe that it is impossible to reduce \( \sigma(x(y).\text{Stop} \uparrow x(z).\text{Stop}) \downarrow R \) into a must-divergent process, since the process becomes successful if one of the components \( \sigma(x(y).\text{Stop}) \) or \( \sigma(x(z).\text{Stop}) \) is part of the redex, and otherwise (the interaction between these two processes is always possible). Thus, \( \sigma(x(y).\text{Stop} \uparrow x(z).\text{Stop}) \downarrow R \) is must-convergent. Hence the preconditions of Theorem D.1 hold and the equivalence holds.
For stop-free processes, contextual equivalence coincides with so-called barbed testing equivalence, where the observation is whether a process may- or should-reduces to a process that has a free input on a fixed channel name (see [24]):

**Definition D.3.** Let $\Pi$ be the subcalculus of $\Pi_{\text{Stop}}$ that does not have the constant $\text{Stop}$ as a syntactic construct. Processes, contexts, reduction, structural congruences are accordingly adapted for $\Pi$.

Let $P \in \Pi$ and $x \in \mathcal{N}$. A process $P$ has a barb on input $x$ (written $P \uparrow x$) iff $P \equiv \nu X.(x(y)P' \downarrow P'')$ where $x \notin X$. We write $P \downarrow x$ iff there exists $P'$ such that $\text{P sr, }^* \longrightarrow P'$ and $P' \downarrow P'$. We write $P \downarrow x$ iff for all $P'$ with $P \text{ sr, }^* \longrightarrow P'$ also $P' \downarrow x$ holds. We write $P \downarrow x$ iff $P \downarrow x$ does not hold, and $P \downarrow x$ iff $P \downarrow x$ does not hold.

For a name $x \in \mathcal{N}$, barbed may- and should-testing preorder $\leq_{c,\text{barb}}$ and barbed may- and should-testing equivalence $\sim_{c,\text{barb}}$ are defined as $\leq_{c,\text{barb}} := \leq_{c,\downarrow} \cap \leq_{c,\downarrow}$ and $\sim_{c,\text{barb}} := \leq_{c,\text{barb}} \cap (\leq_{c,\text{barb}})^{-1}$ where for $\xi \in \{\downarrow, \downarrow'\}$ $\leq_{c,\xi}$ holds iff for all contexts $C \in \Pi : C[P] \xi \implies C[Q] \xi$.

**Theorem D.4 ([24]).** For all processes $P, Q \in \Pi$: $P \leq_{c,\text{barb}} Q \iff P \leq_{c} Q$, and hence also $P \sim_{c,\text{barb}} Q \iff P \sim_{c} Q$. 

Embedding the Pi-Calculus into CH 53