Reducing Weak to Strong Bisimilarity in CCP
Andrés Aristizábal, Filippo Bonchi, Luis Pino, Frank D. Valencia

To cite this version:

HAL Id: hal-00761611
https://hal.archives-ouvertes.fr/hal-00761611
Submitted on 5 Dec 2012

HAL is a multi-disciplinary open access archive for the deposit and dissemination of scientific research documents, whether they are published or not. The documents may come from teaching and research institutions in France or abroad, or from public or private research centers.

L’archive ouverte pluridisciplinaire HAL, est destinée au dépôt et à la diffusion de documents scientifiques de niveau recherche, publiés ou non, émanant des établissements d’enseignement et de recherche français ou étrangers, des laboratoires publics ou privés.
Reducing Weak to Strong Bisimilarity in CCP *

Andrés Aristizábal
CNRS/DGA and LIX École Polytechnique de Paris

Filippo Bonchi
ENS Lyon, Université de Lyon, LIP (UMR 5668 CNRS ENS Lyon UCBL INRIA), 46 Allée d’Italie, 69364 Lyon, France

Luis Pino
Frank Valencia
INRIA/DGA and LIX École Polytechnique de Paris

Concurrent constraint programming (ccp) is a well-established model for concurrency that singles out the fundamental aspects of asynchronous systems whose agents (or processes) evolve by posting and querying (partial) information in a global medium. Bisimilarity is a standard behavioural equivalence in concurrency theory. However, only recently a well-behaved notion of bisimilarity for ccp, and a ccp partition refinement algorithm for deciding the strong version of this equivalence have been proposed. Weak bisimilarity is a central behavioural equivalence in process calculi and it is obtained from the strong case by taking into account only the actions that are observable in the system. Typically, the standard partition refinement can also be used for deciding weak bisimilarity simply by using Milner’s reduction from weak to strong bisimilarity; a technique referred to as saturation. In this paper we demonstrate that, because of its involved labeled transitions, the above-mentioned saturation technique does not work for ccp. We give an alternative reduction from weak ccp bisimilarity to the strong one that allows us to use the ccp partition refinement algorithm for deciding this equivalence.

1 Introduction

Since the introduction of process calculi, one of the richest sources of foundational investigations stemmed from the analysis of behavioural equivalences: in any formal process language, systems which are syntactically different may denote the same process, i.e., they have the same observable behaviour.

A major dichotomy among behavioural equivalences concerns strong and weak equivalences. In strong equivalences, all the transitions performed by a system are deemed observable. In weak equivalences, instead, internal transitions (usually denoted by \( \tau \)) are unobservable. On the one hand, weak equivalences are more abstract (and thus closer to the intuitive notion of behaviour); on the other hand, strong equivalences are usually much easier to be checked (for instance, in [17] a strong equivalence is introduced which is computable for a Turing complete formalism).

Strong bisimilarity is one of the most studied behavioural equivalence and many algorithms (e.g., [30, 11, 12]) have been developed to check whether two systems are equivalent up to strong bisimilarity. Among these, the partition refinement algorithm [15] is one of the best known: first it generates the state

---

*This work has been partially supported by the project ANR-09-BLAN-0169-01 PANDA, and by the French Defence procurement agency (DGA) with two PhD grants.
Reducing Weak to Strong Bisimilarity in CCP

space of a labeled transition system (LTS), i.e., the set of states reachable through the transitions; then, it creates a partition equating all states and afterwards, iteratively, refines these partitions by splitting non equivalent states. At the end, the resulting partition equates all and only bisimilar states.

Weak bisimilarity can be computed by reducing it to strong bisimilarity. Given an LTS $\rightarrow$ labeled with actions $a, b, \ldots$ one can build $\Rightarrow$ as follows.

$$
\begin{align*}
P & \xrightarrow{a} Q \\
P & \Rightarrow Q
\end{align*}
$$

Since weak bisimilarity on $\xrightarrow{a}$ coincides with strong bisimilarity on $\Rightarrow$, then one can check weak bisimilarity with the algorithms for strong bisimilarity on the new LTS $\Rightarrow$.

It is worth pointing out that an alternative presentation of $\Rightarrow$ with sequences of actions as labels is also possible [19]. Nevertheless, the resulting transition system may be infinite-branching and hence not amenable to automatic verification using standard algorithms such as partition refinement.

**Concurrent Constraint Programming** (ccp) [26] is a formalism that combines the traditional algebraic and operational view of process calculi with a declarative one based upon first-order logic. In ccp, processes (agents or programs) interact by adding (or telling) and asking information (namely, constraints) in a medium (the store).

Inspired by [7, 6], the authors introduced in [2] both strong and weak bisimilarity for ccp and showed that the weak equivalence is fully abstract with respect to the standard observational equivalence of [27]. Moreover, a variant of the partition refinement algorithm is given in [3] for checking strong bisimilarity on (the finite fragment) of concurrent constraint programming.

In this paper, first we show that the standard method for reducing weak to strong bisimilarity does not work for ccp and then we provide a way out of the impasse. Our solution can be readily explained by observing that the labels in the LTS of a ccp agent are constraints (actually, they are “the minimal constraints” that the store should satisfy in order to make the agent progress). These constraints form a lattice where the least upper bound (denoted by $\sqcup$) intuitively corresponds to conjunction and the bottom element is the constraint $true$. (As expected, transitions labeled by $true$ are internal transitions, corresponding to the $\tau$ moves in standard process calculi). Now, rather than closing the transitions just with respect to $true$, we need to close them w.r.t. all the constraints. Formally we build the new LTS with the following rules.

$$
\begin{align*}
P & \xrightarrow{a} Q \\
P & \Rightarrow Q
\end{align*}
$$

Note that, since $\sqcup$ is idempotent, if the original LTS $\xrightarrow{a}$ has finitely many transitions, then also $\Rightarrow$ is finite. This allows us to use the algorithm in [3] to check weak bisimilarity on (the finite fragment) of concurrent constraint programming. We have implemented this procedure in a tool that is available at [http://www.lix.polytechnique.fr/~andresaristi/checkers/](http://www.lix.polytechnique.fr/~andresaristi/checkers/) To the best of our knowledge, this is the first tool for checking weak equivalence of ccp programs.

This paper is structured as follows. In Sec. 2 we recall the partition refinement method and the standard reduction from weak to strong bisimilarity. We also recall the ccp formalism, its equivalences, and the ccp partition refinement algorithm. We then show why the standard reduction does not work for ccp. Finally, in Sec. 3 we present our reduction and show its correctness.

**Related Work.** Ccp is not the only formalism where weak bisimilarity cannot be naively reduced to the strong one. Probably the first case in literature can be found in [30] that introduces an algorithm
for checking weak open bisimilarity of $\pi$-calculus. This algorithm is rather different from ours, since it is on-the-fly [11] and thus it checks the equivalence of only two given states (while our algorithm, and more generally all algorithms based on partition refinement, check the equivalence of all the states of a given LTS). Also [4] defines weak labelled transitions following the above-mentioned standard method which does not work in the ccp case.

Analogous problems to the one discussed in this paper arise in Petri nets [28, 10], in tile transition systems [13, 9] and, more generally, in the theory of reactive systems [14] (the interested reader is referred to [29] for an overview). In all these cases, labels form a monoid where the neutral element is the label of internal transitions. Roughly, when reducing from weak to strong bisimilarity, one needs to close the transitions with respect to the composition of the monoid (and not only with respect to the neutral element). However, in all these cases, labels composition is not idempotent (as it is for ccp) and thus a finite LTS might be transformed into an infinite one. For this reason, this procedure applied to the aforementioned cases is not effective for automatic verification.

2 From Weak to Strong CCP Bisimilarity: Saturation Approach

The problem of whether two states are weakly bisimilar in traditional labeled transitions systems is typically reduced to the problem of whether they are strongly bisimilar which can be efficiently verified using partition refinement. We shall refer to this standard reduction as Milner’s saturation method [1].

In this section we shall show that this method does not work for ccp. More precisely, Milner’s reduction will produce an equivalence that does not correspond to the one expected. First, we shall recall the partition refinement algorithm for strong bisimilarity and Milner’s saturation method. Then we show the corresponding notions in ccp.

Standard Partition Refinement. In this section we recall the partition refinement algorithm [15] for checking bisimilarity over the states of a labeled transition system. Remember that an LTS can be intuitively seen as a graph where nodes represent states and arcs represent transitions between states. A transition $P \xrightarrow{a} Q$ between $P$ and $Q$ labeled with $a$ can be typically thought of as an evolution from $P$ to $Q$ provided that a condition $a$ is met. Transition systems can be used to represent the evolution of processes in calculi such as CCS and the $\pi$-calculus [19, 20]. In this case states correspond to processes and transitions are given by the operational semantics of the calculus.

Let us now introduce some notation. Given a set $S$, a partition $\mathcal{P}$ of $S$ is a set of non-empty blocks, i.e., subsets of $S$, that are all disjoint and whose union is $S$. We write $\{B_1\} \ldots \{B_n\}$ to denote a partition consisting of (non-empty) blocks $B_1, \ldots, B_n$. A partition represents an equivalence relation where equivalent elements belong to the same block. We write $P \equiv Q$ to mean that $P$ and $Q$ are equivalent in the partition $\mathcal{P}$.

The partition refinement algorithm (see Alg. [1]) checks bisimilarity as follows. First, it computes $\mathcal{P}^0$, that is the set of all states that are reachable from the set of initial state $\mathcal{I}$. Then it creates the partition $\mathcal{P}^0$ where all the elements of $\mathcal{I}$ belong to the same block (i.e., they are all equivalent). After the initialization, it iteratively refines the partitions by employing the function $\mathcal{F}$, defined as follows: for all partitions $\mathcal{P}$, $P \mathcal{F} Q$ if

- if $P \xrightarrow{a} P'$ then exists $Q'$ s.t. $Q \xrightarrow{a} Q'$ and $P' \mathcal{P} Q'$.

The algorithm terminates whenever two consecutive partitions are equivalent. In such a partition two states belong to the same block iff they are bisimilar.
Algorithm 1 Partition-Refinement (IS)

Initialization

1. $IS^*$ is the set of all processes reachable from $IS$,
2. $P^0 := \{IS^*\}$,

Iteration $P^{n+1} := \mathcal{F}(P^n)$,

Termination If $P^n = P^{n+1}$ then return $P^n$.

<table>
<thead>
<tr>
<th>MR1</th>
<th>$\gamma \xrightarrow{\alpha} \gamma'$</th>
<th>$\gamma \xrightarrow{\alpha} \gamma'$</th>
</tr>
</thead>
<tbody>
<tr>
<td>MR2</td>
<td>$\gamma \xrightarrow{true} \gamma$</td>
<td>$\gamma \xrightarrow{\alpha} \gamma$</td>
</tr>
<tr>
<td>MR3</td>
<td>$\gamma \xrightarrow{true} \gamma \xrightarrow{\alpha} \gamma \xrightarrow{true} \gamma'$</td>
<td>$\gamma \xrightarrow{\alpha} \gamma \xrightarrow{true} \gamma'$</td>
</tr>
</tbody>
</table>

Table 1: Milner’s Saturation Method

Standard reduction from weak to strong bisimilarity. As pointed out in the literature (Chapter 3 from [24]), in order to compute weak bisimilarity, we can use the above mentioned partition refinement. The idea is to start from the graph generated via the operational semantics and then saturate it using the rules described in Tab. 1 to produce a new labeled transition relation $\Longrightarrow$. Recall that $\rightarrow^*$ is the reflexive and transitive closure of the transition relation $\rightarrow$. Now the problem whether two states are weakly bisimilar can be reduced to checking whether they are strongly bisimilar wrt $\Longrightarrow$ using partition refinement. As we will show later on, this approach does not work in a formalism like concurrent constraint programming. We shall see that the problem involves the ccp transition labels which, being constraints, can be arbitrary combined using the lub operation $\sqcup$ to form a new one. Such a situation does not arise in CCS-like labelled transitions.

Notation 1. When the label of a transition is true we will omit it. Namely, henceforth we will use $\gamma \xrightarrow{\alpha} \gamma'$ and $\gamma \Longrightarrow \gamma'$ to denote $\gamma \xrightarrow{true} \gamma'$ and $\gamma \Longrightarrow \gamma'$.

2.1 CCP

We shall now recall ccp and the adaptation of the partition refinement algorithm to compute bisimilarity in ccp [3].

Constraint Systems. The ccp model is parametric in a constraint system (cs) specifying the structure and interdependencies of the information that processes can ask or add to a central shared store. This information is represented as assertions traditionally referred to as constraints. Following [5] [18] we regard a cs as a complete algebraic lattice in which the ordering $\sqsubseteq$ is the reverse of an entailment relation: $c \sqsubseteq d$ means $d$ entails $c$, i.e., $d$ contains “more information” than $c$. The top element false represents inconsistency, the bottom element true is the empty constraint, and the least upper bound (lub) $\sqcup$ is the join of information.

Definition 1 (cs). A constraint system (cs) $C = (Con, Con_0, \sqsubseteq, \sqcup, true, false)$ is a complete algebraic lattice where $Con$, the set of constraints, is a partially ordered set wrt $\sqsubseteq$, $Con_0$ is the subset of compact elements of $Con$, $\sqcup$ is the lub operation defined on all subsets, and true, false are the least and greatest elements of $Con$, respectively.
enough to require that bisimulations are upward closed (iii) they should be equivalent under all the possible contexts of the language. In the case of ccp, it is moves then the other should reply and arrive at an equivalent state (i.e. follow the bisimulation game), two states to be saturated bisimilar, then (i) they should expose the same barbs, (ii) whenever one of them satisfies a weak barb c, (iii) they should be equivalent under all the possible contexts of the language. In the case of ccp, it is enough to require that bisimulations are upward closed as in condition (iii) below.

Remark 1. We shall assume that the constraint system is well-founded and, for practical reasons, that its ordering ⊆ is decidable.

We now define the constraint system we use in our examples.

Example 1. Let Var be a set of variables and ω be the set of natural numbers. A variable assignment is a function μ : Var → ω. We use $\mathcal{A}$ to denote the set of all assignments, $\mathcal{P}(\mathcal{A})$ to denote the powerset of $\mathcal{A}$, $\emptyset$ the empty set and $\cap$ the intersection of sets. Let us define the following constraint system: The set of constraints is $\mathcal{P}(\mathcal{A})$. We define $c \subseteq d$ iff $c \subseteq d$. The constraint false is $\emptyset$, while true is $\mathcal{A}$. Given two constraints $c$ and $d$, $c \cap d$ is the intersection $c \cap d$. We will often use a formula like $x < n$ to denote the corresponding constraint, i.e., the set of all assignments that map $x$ to a number smaller than $n$.

Processes We now recall the basic ccp process constructions. For the sake of space and simplicity we dispense with the recursion operator, which is defined in the standard way as in CCS or other process algebras, and the local/hiding operator (see [2] for further details).

Syntax. Let us presuppose a constraint system $C = (Con, Con_0, \subseteq, \sqcup, true, false)$. The ccp processes are given by the following syntax:

$$P, Q ::= \text{stop} | \text{tell}(c) | \text{ask}(c) \rightarrow P | P \parallel Q | P + Q$$

where $c \in Con_0$. Intuitively, stop represents termination, tell($c$) adds the constraint (or partial information) $c$ to the store. The addition is performed regardless the generation of inconsistent information. The process ask($c$) → $P$ may execute $P$ if $c$ is entailed from the information in the store. The processes $P \parallel Q$ and $P + Q$ stand, respectively, for the parallel execution and non-deterministic choice of $P$ and $Q$.

Reduction Semantics. The operational semantics is given by transitions between configurations. A configuration is a pair $\langle P, d \rangle$ representing a state of a system; $d$ is a constraint representing the global store, and $P$ is a process, i.e., a term of the syntax. We use $Conf$ with typical elements $\gamma, \gamma', \ldots$ to denote the set of configurations. The operational model of ccp is given by the transition relation $\rightarrow$ ⊆ $Conf \times Conf$ defined in Tab. 2. The rules in Tab. 2 are easily seen to realize the above intuitions.

Barbed Semantics. The authors in [2] introduced a barbed semantics for ccp. Barbed equivalences have been introduced in [21] for CCS, and have become the standard behavioural equivalences for formalisms equipped with unlabeled reduction semantics. Intuitively, barbs are basic observations (predicates) on the states of a system. In the case of ccp, bars are taken from the underlying set $Con_0$ of the constraint system. A configuration $\gamma = \langle P, d \rangle$ is said to satisfy the barb $c$ ($\gamma \downarrow c$) iff $c \subseteq d$. Similarly, $\gamma$ satisfies a weak barb $c$ ($\gamma \downarrow c$) iff there exist $\gamma'$ s.t. $\gamma \rightarrow^* \gamma' \downarrow c$.

In this context, the equivalence proposed is the saturated bisimilarity [7, 6]. Intuitively, in order for two states to be saturated bisimilar, then (i) they should expose the same barbs, (ii) whenever one of them moves then the other should reply and arrive at an equivalent state (i.e. follow the bisimulation game), (iii) they should be equivalent under all the possible contexts of the language. In the case of ccp, it is enough to require that bisimulations are upward closed as in condition (iii) below.

Table 2: Reduction semantics for ccp (the symmetric rules for R3 and R4 are omitted).
Reducing Weak to Strong Bisimilarity in CCP

LR1 \(\text{tell}(c), d \xrightarrow{\text{nie}} \text{stop}, d \sqcup c\)

LR2 \(\alpha \in \min\{a \in \text{Con}_0 | c \sqsubseteq d \sqcup a\}\)

LR3 \(\langle P, d \rangle \xrightarrow{\alpha} \langle P', d' \rangle\)

LR4 \(\langle P, d \rangle \xrightarrow{\alpha} \langle P', d' \rangle\)

| Table 3: Labeled semantics for ccp (the symmetric rules for LR3 and LR4 are omitted). |

**Table 3: Labeled semantics for ccp (the symmetric rules for LR3 and LR4 are omitted).**

\[\langle P, d \rangle \xrightarrow{\alpha} \langle P', d' \rangle\]

**Definition 2** (Saturated Barbed Bisimilarity). A saturated barbed bisimulation is a symmetric relation \(R\) on configurations s.t. whenever \((\gamma_1, \gamma_2) \in R\) with \(\gamma_1 = \langle P, c \rangle\) and \(\gamma_2 = \langle Q, d \rangle\) implies that: (i) if \(\gamma_1 \downarrow_c\) then \(\gamma_2 \downarrow_c\), (ii) if \(\gamma_1 \rightarrow \gamma'_1\) then there exists \(\gamma'_2\) s.t. \(\gamma_2 \rightarrow \gamma'_2\) and \((\gamma'_1, \gamma'_2) \in R\), (iii) for every \(a \in \text{Con}_0\), \((\langle P, c \sqcup a \rangle, \langle Q, d \sqcup a \rangle) \in R\). We say that \(\gamma_1\) and \(\gamma_2\) are saturated barbed bisimilar \((\gamma_1 \sim_{sb} \gamma_2)\) if there exists a saturated barbed bisimulation \(R\) s.t. \((\gamma_1, \gamma_2) \in R\).

We use the term “saturated” to be consistent with the original idea in [2]. However, “saturated” in this context has nothing to do with the Milner’s “saturation” for weak bisimilarity. In the following, we will continue to use “saturated” and “saturation” to denote these two different concepts.

**Example 2.** Take \(T = \text{tell}(true)\), \(P = \text{ask}(x < 7) \rightarrow T\) and \(Q = \text{ask}(x < 5) \rightarrow T\). You can see that \(\langle P, true \rangle \not\sim_{sb} \langle Q, true \rangle\), since \(\langle P, x < 7 \rangle \rightarrow\), while \(\langle Q, x < 7 \rangle \rightarrow\). Consider now the configuration \(\langle P + Q, true \rangle\) and observe that \(\langle P + Q, true \rangle \sim_{sb} \langle P, true \rangle\). Indeed, for all constraints \(e\), s.t. \(x < 7 \subseteq e\), both the configurations evolve into \(\langle T, e \rangle\), while for all \(e\), s.t. \(x < 7 \subseteq e\), both configurations cannot proceed. Since \(x < 7 \subseteq x < 5\), the behaviour of \(Q\) is somehow absorbed by the behaviour of \(P\).

As we mentioned before, we are interested in deciding the weak version of the notion above. Then, weak saturated barbed bisimilarity \((\sim_{sb})\) is obtained from Def. [2] by replacing the strong bars in condition (i) for its weak version (\(\not\sim\)) and the transitions in condition (ii) for the reflexive and transitive closure of the transition relation (\(\rightarrow^*\)).

**Labeled Semantics.** As explained in [2], in a transition of the form \(\langle P, d \rangle \xrightarrow{\alpha} \langle P', d' \rangle\) the label \(\alpha\) represents a minimal information (from the environment) that needs to be added to the store \(d\) to evolve from \(\langle P, d \rangle\) into \(\langle P', d' \rangle\), i.e., \(\langle P, d \sqcup \alpha \rangle \rightarrow \langle P', d' \rangle\). The labeled transition relation \(\rightarrow \subseteq \text{Conf} \times \text{Con}_0 \times \text{Conf}\) is defined by the rules in Tab. 3. The rule LR2, for example, says that \(\langle \text{ask}(c), P, d \rangle\) can evolve to \(\langle P, d \sqcup \alpha \rangle\) if the environment provides a minimal constraint \(\alpha\) that added to the store \(d\) entails \(c\), i.e., \(\alpha \in \min\{a \in \text{Con}_0 | c \sqsubseteq d \sqcup a\}\). Note that assuming that \(\text{Con}_0\) is well-founded (Remark [1]) is necessary to guarantee that \(\alpha\) exists whenever \(\{a \in \text{Con}_0 | c \sqsubseteq d \sqcup a\}\) is not empty. The other rules are easily seen to realize the above intuition. Fig. 1 illustrates the LTSs of our running example.

The labeled semantics is sound and complete wrt the unlabeled one. Soundness states that \(\langle P, d \rangle \xrightarrow{\alpha} \langle P', d' \rangle\) corresponds to our intuition that if \(\alpha\) is added to \(d\), \(P\) can reach \(\langle P', d' \rangle\). Completeness states that if we add \(a\) to (the store in) \(\langle P, d \rangle\) and reduce to \(\langle P', d' \rangle\), it exists a minimal information \(\alpha \sqsubseteq a\) such that \(\langle P, d \rangle \xrightarrow{\alpha} \langle P', d'' \rangle\) with \(d'' \sqsubseteq d'\).

The following lemma is an extension of the one in [2] which considers nondeterministic ccpp.

**Lemma 1** (Correctness of \(\rightarrow\)). (Soundness) If \(\langle P, c \rangle \xrightarrow{\alpha} \langle P', c' \rangle\) then \(\langle P, c \sqcup \alpha \rangle \rightarrow \langle P', c' \rangle\). (Completeness) If \(\langle P, c \sqcup a \rangle \rightarrow \langle P', c' \rangle\) then there exists \(\alpha\) and \(b\) s.t. \(\langle P, c \rangle \xrightarrow{\alpha} \langle P', c'' \rangle\) where \(\alpha \sqcap b = a\) and \(c' \sqcup b = c'\).

The above lemma is central for deciding bisimilarity in ccpp. In fact, we will show later that for the weak (saturated) semantics the completeness direction does not hold. From this we will show that the standard reduction from weak to strong does not work.
Transition (a) means that for all constraints \( (c) \) is possible for all its meaning is “logically derived” by transition (a). The following notion captures the above intuition:

\[
\begin{align*}
(T + S, \text{true}) & \xrightarrow{\langle z < 5 \rangle} \langle P + Q', z < 5 \rangle \quad \text{true} \quad \langle \text{stop}, z < 5 \iff x < 5 \land y = 1 \rangle \\
(S, \text{true}) & \xrightarrow{\langle z < 7 \rangle} \langle P, z < 7 \rangle \quad \text{true} \quad \langle \text{stop}, z < 7 \iff x < 7 \rangle \\
(R + S, \text{true}) & \xrightarrow{\langle z < 5 \rangle} \langle P + Q, z < 5 \rangle \quad \text{true} \quad \langle \text{stop}, z < 5 \iff x < 5 \rangle \\
\end{align*}
\]

Figure 1: The LTS of the running example \((IS = \{ \langle R' + S, \text{true} \rangle, \langle S, \text{true} \rangle, \langle R + S, \text{true} \rangle \})\).

### 2.1.1 Equivalences: Saturated Barbed, Irredundant and Symbolic Bisimilarity

In this section we recall how to check \( \sim_{sb} \) with a modified version of partition refinement introduced in \([3]\). Henceforth, we shall refer to this version as \( \text{ccp partition refinement} \) (ccp-PR).

The main problem with checking \( \sim_{sb} \) is the quantification over all contexts. This problem is addressed in \([3]\) following the abstract approach in \([8]\). More precisely, we use an equivalent notion, namely \( \text{irredundant bisimilarity} \sim_{I} \), which can be verified with ccp-PR. As its name suggests, \( \sim_{I} \) only takes into account those transitions deemed irredundant.\(^1\) However, technically speaking, going from \( \sim_{sb} \) to \( \sim_{I} \) requires one intermediate notion, so-called \( \text{symbolic bisimilarity} \). These three notions are shown to be equivalent, i.e., \( \sim_{sb} \equiv \sim_{sym} \equiv \sim_{I} \). In the following we recall all of them.

Let us first give some auxiliary definitions. The first concept is that of \textit{derivation}. Consider the following transitions (taken from Fig. [1]):

\[
\begin{align*}
\text{(a)} \quad \langle P + Q, z < 5 \rangle & \xrightarrow{x < 7} \langle T, z < 5 \land x < 7 \rangle & \quad \text{(b)} \quad \langle P + Q, z < 5 \rangle & \xrightarrow{x < 5} \langle T, z < 5 \land x < 5 \rangle
\end{align*}
\]

Transition (a) means that for all constraints \( e \) s.t. \( x < 7 \) is entailed by \( e \) (formally \( x < 7 \sqsubseteq e \)), the transition (c) \( \langle P + Q, z < 5 \land e \rangle \rightarrow \langle T, z < 5 \land e \rangle \) can be performed, while transition (b) means that the reduction (c) is possible for all \( e \) s.t. \( x < 5 \sqsubseteq e \). Since \( x < 7 \sqsubseteq x < 5 \), transition (b) is “redundant”, in the sense that its meaning is “logically derived” by transition (a). The following notion captures the above intuition:

\[\text{Definition 3 (Derivation } \vdash_D \text{). We say that the transition } t = \langle P, c \rangle \xrightarrow{\alpha} \langle P', c' \rangle \text{ derives } t' = \langle P, c \rangle \xrightarrow{\beta} \langle P', c'' \rangle \text{ (written } t \vdash_D t' \text{) iff there exists } e \text{ s.t. } \alpha \sqcap e = \beta \text{ and } c' \sqcap e = c''.\]

One can verify in the above example that (a) \( \vdash_D \) (b), and notice that both transitions arrive at the same process \( P' \), the difference lies in the label and the store. Now imagine the situation where the initial configuration is able to perform another transition with \( \beta \) (as in \( t' \)), let us also assume that such transition arrives at a configuration which is equivalent to the result of \( t' \). Therefore, it is natural to think that, since \( t \) dominates \( t' \), such new transition should also be dominated by \( t \). Let us explain with an example, consider the two following transitions:

\[
\begin{align*}
\text{(e)} \quad \langle R + S, \text{true} \rangle & \xrightarrow{z < 7} \langle P, z < 7 \rangle & \quad \text{(f)} \quad \langle R + S, \text{true} \rangle & \xrightarrow{z < 5} \langle P + Q, z < 5 \rangle
\end{align*}
\]

\(^1\)Redundancy itself is not trivial to check, for more information go to \([3]\).
Note that transition (f) cannot be derived by other transitions, since (e) \( \not\vdash_D \) (f). Indeed, \( P \) is syntactically different from \( P + Q \), even if they have the same behaviour when inserted in the store \( z < 5 \), i.e., \( (P, z < 5) \sim_{sb} (P + Q, z < 5) \) (since \( \sim_{sb} \) is upward closed). Transition (f) is also “redundant”, since its behaviour “does not add anything” to the behavior of (e). The following definition encompasses this situation:

**Definition 4** (Derivation w.r.t. \( R, \vdash_R \)). We say that the transition \( t = \gamma \xrightarrow{\alpha} \gamma_1 \) derives \( t' = \gamma \xrightarrow{\beta} \gamma_2 \) w.r.t. \( R \) (written \( t \vdash_R t' \)) iff there exists \( \gamma_2 \) s.t. \( t \vdash_D \gamma \xrightarrow{\beta} \gamma_2 \) and \( \gamma_2 \vdash_R \gamma_2' \).

Then, when \( R \) represents some sort of equivalence, this notion will capture the situation above mentioned. Notice that \( \vdash_D \) is \( \vdash_R \) with \( R \) being the identity relation (id). Now we introduce the concept of domination, which consists in strengthening the notion of derivation by requiring labels to be different.

**Definition 5** (Domination \( \succ_D \)). We say that the transition \( t = (P, c) \xrightarrow{\alpha} (P', c') \) dominates \( t' = (P, c) \xrightarrow{\beta} (P', c'') \) (written \( t \succ_D t' \)) if \( t \vdash_D (P, c) \xrightarrow{\beta} (P', c'') \) and \( \alpha \neq \beta \).

Similarly, as we did for derivation, we can define domination depending on a relation. Again, \( \succ_D \) is just \( \succ_R \) when \( R \) is the identity relation (id).

**Definition 6** (Redundancy and Domination w.r.t. \( R, \succ_R \)). We say that the transition \( t = (P, c) \xrightarrow{\alpha} (P', c') \) dominates \( t' = (P, c) \xrightarrow{\beta} (P', c'') \) w.r.t. \( R \) (written \( t \succ_R t' \)) iff there exists \( c'' \) s.t. \( t \succ_D (P, c) \xrightarrow{\beta} (P', c'') \) and \( (P, c''), \succ_R (Q, d) \). Also, a transition is said to be redundant when it is dominated by another, otherwise it is said to be irredundant.

We are now able to introduce symbolic bisimilarity. Intuitively, two configurations \( \gamma_1 \) and \( \gamma_2 \) are symbolic bisimilar iff (i) they have the same barbs and (ii) whenever there is a transition from \( \gamma_1 \) to \( \gamma_1' \) using \( \alpha \), then we require that \( \gamma_2 \) must reply with a similar transition \( \gamma_2 \xrightarrow{\alpha} \gamma_2' \) (where \( \gamma_1' \) and \( \gamma_2' \) are now equivalent) or some other transition that derives it. In other words, the move from the defender does not need to use exactly the same label, but a transition that is “stronger” (in terms of derivation \( \vdash_D \)) could also do the job. Formally we have the definition below.

**Definition 7** (Symbolic Bisimilarity). A symbolic bisimulation is a symmetric relation \( \mathcal{R} \) on configurations s.t. whenever \( (\gamma_1, \gamma_2) \in \mathcal{R} \) with \( \gamma_1 = (P, c) \) and \( \gamma_2 = (Q, d) \) implies that: (i) if \( \gamma_2 \Downarrow_e \) then \( \gamma_2 \Downarrow_e \), (ii) if \( (P, c) \xrightarrow{\alpha} (P', c') \) then there exists a transition \( t = (Q, d) \xrightarrow{\beta} (Q', d') \) and a store \( d' \) s.t. \( t \vdash_D (Q, d) \xrightarrow{\alpha} (Q', d') \) and \( (P, c'), \succ_R (Q', d') \). We say that \( \gamma_1 \) and \( \gamma_2 \) are symbolic bisimilar (\( \gamma_1 \sim_{sb} \gamma_2 \)) if there exists a symbolic bisimulation \( \mathcal{R} \) s.t. \( (\gamma_1, \gamma_2) \in \mathcal{R} \).

**Example 3.** To illustrate the notion of \( \sim_{sb} \) we take \( (P + Q, \text{true}) \) and \( (P, \text{true}) \) from Ex. 2. We provide a symbolic bisimulation \( \mathcal{R} = \{ ((P + Q, \text{true}), (P, \text{true})) \} \cup \text{id} \) to prove \( (P + Q, \text{true}) \sim_{sb} (P, \text{true}) \). We take the pair \( ((P + Q, \text{true}), (P, \text{true})) \). The first condition in Def. 7 is trivial. For the second one, we take \( (P + Q, \text{true}) \xrightarrow{\text{inc5}} (T, x < 5) \) and one can find transitions \( t = (P, \text{true}) \xrightarrow{\text{inc5}} (T, x < 5) \) and \( t' = (P, \text{true}) \xrightarrow{\text{inc5}} (T, x < 5) \) s.t. \( t \vdash_D t' \) and \( (T, x < 5) \mathcal{R} (T, x < 5) \). The restant pairs are trivially verified.

And finally, the irredundant version, which follows the standard bisimulation game where labels need to be matched, however only those transitions so-called irredundant must be considered.

**Definition 8** (Irredundant Bisimilarity). An irredundant bisimulation is a symmetric relation \( \mathcal{R} \) on configurations s.t. whenever \( (\gamma_1, \gamma_2) \in \mathcal{R} \) implies that: (i) if \( \gamma_1 \Downarrow_e \) then \( \gamma_2 \Downarrow_e \), (ii) if \( \gamma_1 \xrightarrow{\alpha} \gamma_1' \) and it is irredundant in \( \mathcal{R} \) then there exists \( \gamma_2' \) s.t. \( \gamma_2 \xrightarrow{\alpha} \gamma_2' \) and \( (\gamma_1', \gamma_2') \in \mathcal{R} \). We say that \( \gamma_1 \) and \( \gamma_2 \) are irredundant bisimilar (\( \gamma_1 \sim_{ib} \gamma_2 \)) if there exists an irredundant bisimulation \( \mathcal{R} \) s.t. \( (\gamma_1, \gamma_2) \in \mathcal{R} \).
Example 4. We can verify that the relation $\mathcal{R}$ in Ex. 3 is an irredundant bisimulation to show that $(P + Q, true) \sim (P, true)$. We take the pair $(\langle P + Q, true \rangle, \langle P, true \rangle)$. The first item in Def. 3 is obvious. Then take $(P + Q, true) \xrightarrow{x < 7} \langle T, x < 7 \rangle$, which is irredundant according to Def. 6 and then there exists a $(\langle T, x < 7 \rangle, \langle T, x < 7 \rangle) \in \mathcal{R}$. The other pairs are trivially proven. Notice that $(P + Q, true) \xrightarrow{x < 7} \langle T, x < 7 \rangle \Rightarrow (P + Q, true) \xrightarrow{x < 5} \langle T, x < 5 \rangle$ hence $(P + Q, true) \xrightarrow{x < 5} \langle T, x < 5 \rangle$ is redundant, thus it does not need to be matched by $(P, true)$.

As we said at the beginning, the above-defined equivalences coincide with $\sim_{sb}$. The proof, given in [3], strongly relies on Lemma 1.

Theorem 1. $(P, c) \sim_{1} (Q, d)$ iff $(P, c) \sim_{sym} (Q, d)$ iff $(P, c) \sim_{sb} (Q, d)$

2.1.2 Partition Refinement for CCP

In [3] the authors introduced an algorithm for checking $\sim_{sb}$, by modifying the partition refinement algorithm so that to exploit $\sim_{1}$. First, since configurations satisfying different bars are surely different, it can be safely started with a partition that equates all and only those states satisfying the same bars. Note that two configurations satisfy the same bars iff they have the same store. Thus, we take as initial partition $P^0 = \{IS^{*}_{d_{1}}\} \ldots \{IS^{*}_{d_{n}}\}$, where $IS^{*}_{d}$ is the subset of the configurations of $IS^{*}$ with store $d$.

Secondly, instead of using the function $F$ of Alg. 1, the partitions are refined by employing the function $IR$ defined as follows: for all partitions $P$, $\gamma_1 IR(P) \gamma_2$ if

- if $\gamma_1 \xrightarrow{\alpha} \gamma'_1$ is irredundant in $P$, then there exists $\gamma'_2$ s.t. $\gamma_2 \xrightarrow{\alpha} \gamma'_2$ and $\gamma'_1 P \gamma'_2$.

These two steps are the main idea behind the computation of $\sim_{1}$ (Alg. 2).

Algorithm 2 CCP-Partition-Refinement (IS)

Initialization

1. Compute $IS^{*}_{new}$
2. $P^0 := \{IS^{*}_{d_{1}}\} \ldots \{IS^{*}_{d_{n}}\}$

Iteration $P^{n+1} := IR(P^n)$

Termination If $P^n = P^{n+1}$ then return $P^n$.

2.2 Incompleteness of Milner’s saturation method in ccp

As mentioned at the beginning of this section, the standard approach for deciding weak equivalences is to add some transitions to the original processes, so-called saturation, and then check for the strong equivalence. In calculi like CCS, such saturation consists in forgetting about the internal actions that make part of a sequence containing one observable action (Tab. 1). However, for ccp this method does not work. The problem is that the transition relation proposed by Milner is not complete for ccp, hence the relation among the saturated, symbolic and irredundant equivalences is broken. In the next section we will provide a stronger saturation, which is complete, and allow us to use the ccp-PR to compute $\approx_{sb}$.

Let us show why Milner’s approach does not work. First, we need to introduce formally the concept of completeness for a given transition relation.

2In fact, in order to check redundancy, some new states should be added to the initial ones (hence the subscript new in $IS_{new}$). The details of the computation are omitted given that they are not relevant for this paper, however the interested reader is referred to [3] for more information.
\[\langle \text{ask } \alpha \rightarrow \langle \text{ask } \beta \rightarrow \text{stop}, \alpha \sqcup \beta \rangle \quad \text{and} \quad \langle \text{ask } \alpha \rightarrow \langle \text{ask } \beta \rightarrow \text{stop}, \text{true} \rangle,\beta \rangle\]

Figure 2: Counterexample for completeness using Milner’s saturation method (cycles from MR2 omitted). Both graphs are obtained by applying the rules in Tab. 1.

**Definition 9.** We say that a transition relation \(\sim\subseteq \text{Conf} \times \text{Con}_0 \times \text{Conf}\) is complete iff whenever \(\langle P, c \sqcup a \rangle \sim \langle P', c' \rangle\) then there exist \(\alpha, b \in \text{Con}_0\) s.t. \(\langle P, c \rangle \sim \langle P', c' \rangle\) where \(\alpha \sqcup b = a\) and \(c'' \sqcup b = c'\).

Notice that \(\rightarrow\) (i.e. the reduction semantics, see Table 2) is complete, and it corresponds to the second item of Lemma 1. Now Milner’s method defines a new transition relation \(\Longrightarrow\) using the rules in Tab. 1 but it turns out not to be complete.

**Proposition 1.** The relation \(\Longrightarrow\) defined in Table 7 is not complete.

**Proof.** We will show a counter-example where the completeness for \(\Longrightarrow\) does not hold. Let \(P = \langle \text{ask } \alpha \rightarrow \langle \text{ask } \beta \rightarrow \text{stop} \rangle\) and \(d = \alpha \sqcup \beta\). Now consider the transition \(\langle P, d \rangle \Longrightarrow \langle \text{stop}, d \rangle\) and let us apply the completeness lemma, we can take \(c = \text{true}\) and \(\alpha = \alpha \sqcup \beta\), therefore by completeness there must exist \(b\) and \(\lambda\) s.t. \(\langle P, \text{true} \rangle \xrightarrow{\lambda} \langle \text{stop}, c'' \rangle\) where \(\alpha \sqcup b = \alpha \sqcup \beta\) and \(c'' \sqcup b = d\). However, notice that the only transition possible is \(\langle P, \text{true} \rangle \xrightarrow{\alpha} \langle \text{ask } \beta \rightarrow \text{stop}, \alpha \rangle\), hence completeness does not hold since there is no transition from \(\langle P, \text{true} \rangle\) to \(\langle \text{stop}, c'' \rangle\) for some \(c''\). Fig. 2 illustrates the problem.

We can now use this fact to see why the method does not work for computing \(\approx_{sb}\) using ccp-PR. First, let us redefine some concepts using the new transition relation \(\Longrightarrow\). Because of condition (i) in \(\approx_{sb}\), we need a new definition of bars, namely \emph{weak bars} w.r.t. \(\Longrightarrow\).

**Definition 10.** We say \(\gamma\) has a weak barb \(e\) w.r.t. \(\Longrightarrow\) (written \(\gamma \sqsubseteq_{\text{wbar}} e\)) iff \(\gamma \Longrightarrow^* \gamma \downarrow e\).

Using this notion, we introduce Symbolic and Irredundant bisimilarity w.r.t. \(\Longrightarrow\), denoted by \(\sim_{\text{sym}}\) and \(\sim_{\text{ir}}\) respectively. They are defined as in Def. 7 and 8 where in condition (i) weak bars \(\downarrow\) are replaced with \(\downarrow\) and in condition (ii) the transition relation is now \(\Longrightarrow\).

One would expect that since \(\sim_{sb} = \sim_{\text{sym}} = \sim_{\text{ir}}\) then the natural consequence will be that \(\approx_{sb} = \approx_{\text{sym}} = \approx_{\text{ir}}\), given that these new notions are supposed to be the weak versions of the former ones when using the saturation method. However, completeness is necessary for proving \(\sim_{sb} = \sim_{\text{sym}} = \sim_{\text{ir}}\), and from Proposition 7 we know that \(\Longrightarrow\) is not complete hence we might expect \(\approx_{sb} \neq \approx_{\text{sym}} \neq \approx_{\text{ir}}\). In fact, the following counter-example shows these inequalities.

**Example 5.** Let \(P, P'\) and \(Q\) as in Fig. 2. The figure shows \(\langle P, \text{true} \rangle\) and \(\langle Q, \text{true} \rangle\) after we saturate them using Milner’s method. First, notice that \(\langle P, \text{true} \rangle \approx_{sb} \langle Q, \text{true} \rangle\), since there exists a saturated weak barbed bisimulation \(\mathcal{B} = \{\langle \langle P, \text{true} \rangle, \langle Q, \text{true} \rangle \rangle\} \cup \text{id}\). However, \(\langle P, \text{true} \rangle \not\approx_{\text{ir}} \langle Q, \text{true} \rangle\). To prove that, we need to pick an irredundant transition from \(\langle P, \text{true} \rangle\) or \(\langle Q, \text{true} \rangle\) (after saturation) s.t. the other cannot match. Thus, take \(\langle Q, \text{true} \rangle \xrightarrow{\alpha \sqcup \beta} \langle \text{tell}(c), \alpha \sqcup \beta \rangle\) which is irredundant and given that \(\langle P, \text{true} \rangle\) does not have a transition with \(\alpha \sqcup \beta\) then we know that there is no irredundant bisimulation containing \(\langle \langle P, \text{true} \rangle, \langle Q, \text{true} \rangle \rangle\) therefore \(\langle P, \text{true} \rangle \not\approx_{\text{ir}} \langle Q, \text{true} \rangle\). Using the same reasoning we can also show that \(\approx_{sb} \neq \approx_{\text{sym}}\).
3 Reducing weak bisimilarity to Strong in CCP

In this section we shall provide a method for deciding weak bisimilarity in ccp. As shown in Sec. 2.2, the usual method for deciding weak bisimilarity (introduced in Sec. 2) does not work for ccp. We shall proceed by redefining $\Rightarrow$ in such a way that it is sound and complete for ccp. Then we prove that, w.r.t. $\Rightarrow$, symbolic and irredundant bisimilarity coincide with $\hat{\approx}_{sb}$, i.e. $\approx_{sb} = \hat{\approx}_{sym} = \hat{\approx}_{I}$. We therefore conclude that the partition refinement algorithm in [3] can be used to verify $\approx_{sb}$ w.r.t. $\Rightarrow$.

3.1 Defining a new saturation method for CCP

If we analyze the counter-example to completeness (see Fig. 2), one can see that the problem arises because of the nature of the labels in ccp, namely using this method $\langle \text{ask } \alpha \rightarrow (\text{ask } \beta \rightarrow \text{stop} \rangle$, true) does not have a transition with $\alpha \sqcup \beta$ to $\langle \text{stop} \rangle$, $\alpha \sqcup \beta$, hence that fact can be exploited to break the relation among the weak equivalences. Following this reasoning, instead of only forgetting about the silent actions we also take into account that labels in ccp can be added together. Thus we have a new rule that creates a new transition for each two consecutive ones, whose label is the lub of the labels in them. This method can also be thought as the reflexive and transitive closure of the labeled transition relation $(\rightarrow)$. This transition relation turns out to be sound and complete and it can be used to decide $\approx_{sb}$.
3.1.1 A new saturation method

Formally, our new transition relation $\Rightarrow$ is defined by the rules in Tab. 4. For simplicity, we are using the same arrow $\Rightarrow$ to denote this transition relation. Consequently the definitions of weak barbs, symbolic and redundant bisimilarity are now interpreted w.r.t. $\Rightarrow$ ($\downarrow$, $\downarrow_{sym}$ and $\downarrow_{I}$ respectively).

First, $\downarrow$ coincides with $\downarrow$, since a transition in $\Rightarrow$ corresponds to a sequence of reductions.

Lemma 2. $\gamma \rightarrow^* \gamma'$ iff $\gamma \Rightarrow \gamma'$.

Using this lemma, it is straightforward to see that the notions of weak barbs coincide.

Proposition 2. $\downarrow \downarrow$ iff $\downarrow \beta \downarrow$.

An important property is that the new labeled transition system ($\Rightarrow$) is finitely branching. Under the assumption that the transition relation $\rightarrow$ is finitely branching and that the amount of states in the transition system is finite, this way, we can use the fact that labels in ccp are idempotent to prove that $\Rightarrow$ is finitely branching. Formally:

Proposition 3. If for any $\gamma$ we have $|(\gamma', \alpha)| \gamma \rightarrow \gamma'| < \infty$ and $|(\forall \alpha_1, \ldots, \alpha_n. \gamma \rightarrow \alpha_1 \ldots \alpha_n \rightarrow \gamma')| < \infty$, then $|(\forall \alpha. \gamma \Rightarrow \gamma')| < \infty$.

3.1.2 Soundness and Completeness

As mentioned before, soundness and completeness of the relation are the core properties when proving $\sim_{sb} = \sim_{sym} = \sim_I$. We now proceed to show that our method enjoys of these properties and they will allow us to prove the correspondence among the equivalences for the weak case.

Lemma 3 (Soundness of $\Rightarrow$). If $\langle P, c \rangle \xRightarrow{\alpha} \langle P', c' \rangle$ then $\langle P, c \downarrow \alpha \rangle \Rightarrow \langle P', c' \rangle$.

Proof. We proceed by induction on the depth of the inference of $\langle P, c \rangle \xRightarrow{\alpha} \langle P', c' \rangle$.

- Using R-Tau we have $\langle P, c \rangle \Rightarrow \langle P, c \rangle$ and the result follows directly given that $\alpha = \text{true}$.
- Using R-Label we have $\langle P, c \rangle \xRightarrow{\alpha} \langle P', c' \rangle$ then $\langle P, c \xrightarrow{\alpha} \rangle \Rightarrow \langle P', c' \rangle$. By Lemma 1 (soundness of $\rightarrow$) we get $\langle P, c \downarrow \alpha \rangle \Rightarrow \langle P', c' \rangle$ and finally by rule R-Label $\langle P, c \downarrow \alpha \rangle \Rightarrow \langle P', c' \rangle$.
- Using R-Add then we have $\langle P, c \rangle \xRightarrow{\beta \uparrow \lambda} \langle P', c' \rangle$ then $\langle P, c \rangle \xRightarrow{\beta} \langle P'', c'' \rangle \xRightarrow{\lambda} \langle P', c' \rangle$ where $\beta \uparrow \lambda = \alpha$. By induction hypothesis, $\langle P, c \uparrow \beta \downarrow \lambda \rangle \Rightarrow \langle P'', c'' \rangle (1)$ and $\langle P'', c'' \uparrow \lambda \rangle \Rightarrow \langle P', c' \rangle (2)$. By monotonicity on (1), $\langle P, c \uparrow \beta \downarrow \lambda \rangle \Rightarrow \langle P'', c'' \uparrow \lambda \rangle$ and by rule R-Add on this transition and (2) then, given that $\beta \uparrow \lambda = \alpha$, we obtain $\langle P, c \downarrow \alpha \rangle \Rightarrow \langle P', c' \rangle$.

Lemma 4 (Completeness of $\Rightarrow$). If $\langle P, c \downarrow \alpha \rangle \Rightarrow \langle P', c' \rangle$ then there exist $\alpha$ and $b$ s.t. $\langle P, c \rangle \xRightarrow{\alpha} \langle P', c'' \rangle$ where $\alpha \downarrow b = a$ and $c'' \downarrow b = c'$. 

<table>
<thead>
<tr>
<th>R-Tau</th>
<th>R-Label</th>
<th>R-Add</th>
</tr>
</thead>
<tbody>
<tr>
<td>$\gamma \Rightarrow \gamma$</td>
<td>$\gamma \xrightarrow{\alpha} \gamma'$</td>
<td>$\gamma \xrightarrow{\alpha} \gamma' \xrightarrow{\beta} \gamma''$</td>
</tr>
</tbody>
</table>
Proof. Assuming that \( \langle P, c \sqcap a \rangle \Rightarrow \langle P', c' \rangle \) then, from Lemma 2 we can say that \( \langle P, c \sqcap a \rangle \rightarrow^* \langle P', c' \rangle \) which can be written as \( \langle P, c \sqcap a \rangle \rightarrow \ldots \rightarrow \langle P_i, c_i \rangle \rightarrow \langle P', c' \rangle \), we will proceed by induction on \( i \).

(Base Case) Assuming \( i = 0 \) then \( \langle P, c \sqcap a \rangle \rightarrow \langle P', c' \rangle \) and the result follows directly from Lemma 1 (Completeness of \( \Rightarrow \)) and R-Label.

(Induction) Let us assume that \( \langle P, c \sqcap a \rangle \rightarrow^i \langle P_i, c_i \rangle \rightarrow \langle P', c' \rangle \) then by induction hypothesis there exist \( \beta \) and \( \beta' \) s.t. \( \langle P, c \rangle \xrightarrow{\beta} \langle P_i, c_i \rangle \) (1) where \( \beta \sqcup \beta' = a \) and \( c_i \sqcup b' = c_i \). Now by completeness on the last transition \( \langle P_i, c_i \rangle \rightarrow \langle P', c' \rangle \), there exists \( \lambda \) and \( b'' \) s.t. \( \langle P_i, c_i \rangle \xrightarrow{\lambda} \langle P', c'' \rangle \) where \( \lambda \sqcup b'' = b' \) and \( c'' \sqcup b'' = c' \), thus by rule R-Label we have \( \langle P_i, c_i \rangle \xrightarrow{\beta} \langle P', c'' \rangle \) (2). We can now proceed to apply rule R-Add on (1) and (2) to obtain the transition \( \langle P, c \rangle \xrightarrow{\alpha} \langle P', c'' \rangle \) where \( \alpha = \beta \sqcup \lambda \) and finally take \( b = b'' \), therefore the conditions hold \( \beta \sqcup b = \beta \sqcup \lambda \sqcup b'' = a \) and \( c'' \sqcup b = c'' \sqcup b'' = c' \).

3.2 Weak saturated bisimilarity coincides with the strong symbolic and irredundant bisimilarity

We show our main result, a method for deciding \( \approx \), which has an effective decision procedure. Here we prove that \( \approx \) is sound and complete, the correspondence between the symbolic and irredundant bisimilarity follows from 3.

**Corollary 1.** \( \gamma \approx \gamma ' \) iff \( \gamma \approx \gamma ' \)

Finally, in the next two lemmata, we prove that \( \approx = \approx_{\text{sym}} \).

**Lemma 5.** If \( \gamma \approx \gamma ' \) then \( \gamma \approx \gamma ' \)

**Proof.** We need to prove that \( \mathcal{R} = \{ \langle \langle P, c \rangle, \langle Q, d \rangle \rangle \mid \langle P, c \rangle \approx_{\text{sym}} \langle Q, d \rangle \} \) is a symbolic bisimulation over \( \Longrightarrow \). The first condition (i) of the bisimulation follows directly from Proposition 2. As for (ii), let us assume that \( \langle P, c \rangle \xrightarrow{\alpha} \langle P', c' \rangle \) then by soundness of \( \Rightarrow \) we have \( \langle P, c \sqcup \alpha \rangle \rightarrow \langle P', c' \rangle \), now by Lemma 2 we obtain \( \langle P, c \sqcup \alpha \rangle \rightarrow \langle P', c' \rangle \). Given that \( \langle P, c \rangle \approx_{\text{sym}} \langle Q, d \rangle \) then from the latter transition we can conclude that \( \langle Q, d \sqcup \alpha \rangle \rightarrow \langle Q', d' \rangle \) where \( \langle Q', d' \rangle \approx_{\text{sym}} \langle Q', d' \rangle \), hence we can use Lemma 2 again to deduce that \( \langle Q, d \sqcup \alpha \rangle \rightarrow \langle Q', d' \rangle \). Finally, by completeness of \( \Rightarrow \), there exist \( \beta \) and \( b \) s.t. \( t = \langle Q, d \rangle \xrightarrow{\beta} \langle Q', d'' \rangle \) where \( \beta \sqcup b = \alpha \) and \( d'' \sqcup b = d' \), therefore \( t \vdash_{\mathcal{D}} \langle Q, d \rangle \xrightarrow{\alpha} \langle Q', d'' \rangle \) and \( \langle P', c' \rangle \approx_{\text{sym}} \langle Q', d'' \rangle \).

**Lemma 6.** If \( \gamma \approx \gamma ' \) then \( \gamma \approx_{\text{sym}} \gamma ' \)

**Proof.** We need to prove that \( \mathcal{R} = \{ \langle \langle P, c \rangle, \langle Q, d \rangle \rangle \mid \langle P, c \rangle \approx_{\text{sym}} \langle Q, d \rangle \} \) is a weak saturated bisimulation. First, condition (i) follows form Proposition 2 and (iii) by definition of \( \mathcal{R} \). Let us prove condition (ii), assume \( \langle P, c \sqcup a \rangle \rightarrow \langle P', c' \rangle \) then by Lemma 2 \( \langle P, c \sqcup a \rangle \rightarrow \langle P', c' \rangle \). Now by completeness of \( \Rightarrow \) there exist \( \alpha \) and \( b \) s.t. \( \langle P, c \rangle \xrightarrow{\alpha} \langle P', c'' \rangle \) where \( \alpha \sqcup b = a \) and \( c'' \sqcup b = c' \). Since \( \langle P, c \rangle \approx_{\text{sym}} \langle Q, d \rangle \) then we know there exists a transition \( t = \langle Q, d \rangle \xrightarrow{\beta} \langle Q', d'' \rangle \) s.t. \( t \vdash_{\mathcal{D}} \langle Q, d \rangle \xrightarrow{\alpha} \langle Q', d'' \rangle \).
and (a) \( (P', c'') \sim_{\text{sym}} (Q', d'') \), by definition of \( \vdash D \) there exists \( b' \) s.t. \( \beta \uplus b' = \alpha \) and \( d' \uplus b' = d'' \). Using soundness of \( \Rightarrow \) on \( t \) we get \( (Q, d \uplus \beta) \Rightarrow (Q', d') \), thus by Lemma 2 \( (Q, d \uplus \beta) \rightarrow^* (Q', d') \) and finally by monotonicity

\[
\langle Q, d \uplus \beta \uplus b' \uplus \beta \rangle \rightarrow^* \langle Q', d' \uplus b' \uplus \beta \rangle
\]

(1)

Then, the transition \( \langle P, c \uplus a \rangle \rightarrow^* (P', c') \) can be rewritten as \( \langle P, c \uplus a \rangle \rightarrow^* (P', c'' \uplus b) \), and using (1), \( (Q, d \uplus a) \rightarrow^* (Q', d'' \uplus b) \). It is left to prove that \( \langle P', c'' \uplus b \rangle \triangleright (Q', d'' \uplus b) \) which follows from (a).

4 Concluding Remarks

We showed that the transition relation given by Milner’s saturation method is not complete for ccp (in the sense of Definition 9). As consequence we also showed that weak saturated barbed bisimilarity \( \approx_{sb} \) cannot be computed using the ccp partition refinement algorithm for (strong) bisimilarity ccp wrt to this transition relation. We then presented a new transition relation using another saturation mechanism and showed that it is complete for ccp. We also showed that the ccp partition refinement can be used to compute \( \approx_{sb} \) using the new transition relation. To the best of our knowledge, this is the first approach to verifying weak bisimilarity for ccp. As future work, we plan to investigate other calculi where the nature of their transition systems give rise to similar situations regarding weak and strong bisimilarity, in particular timed ccp (tcc) [25], non-deterministic timed ccp (ntcc) [23], universal temporal ccp (utcc) [22] and Epistemic ccp (ecpp) [16].

References


