ITL Semantics of Composite Petri Nets

Z. Duan, H. Klaudel and M. Koutny

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Suggested keywords

ITL
PETRI NET
BOX ALGEBRA
COMPOSITION
SEMANTICS
ITL Semantics of Composite Petri Nets

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Keywords: \textit{itl}, Petri net, box algebra, composition, semantics.

1 Introduction

Temporal logics \cite{8,3} and Petri nets \cite{13} are two different but, in many yet com-
plementary formalisms for the specification and analysis of concurrent
systems. A temporal logic, such as Interval Temporal Logic (\textit{itl}) \cite{10,12}, allows one to
specify both the system design and correctness requirements within the same
logic framework based on sequences of global states. As a result, verification of
a requirement captured by logic formula $\phi$ for a concurrent system expressed
as logic formula $\psi$ can be done by checking that the implication $\psi \supset \phi$ holds
true. Petri nets, on the other hand, which are a graphical model with semantics
based on local states allow, e.g., for a direct expression of causality aspects in
system behavior. As a result, verification of system properties can be done using
model checking techniques based on partial order reductions \cite{18}, or invariant

techniques \cite{17} based on graph structure of nets.

In this paper, we aim at providing a basic semantical link between \textit{itl} and
Petri nets so that one would be able to use both kinds of verification techniques
of system properties. Finding such a link is bound to be difficult as temporal logics and Petri nets have strikingly different nature. The first step, therefore, would be an identification of a sufficiently expressive temporal logic and class of Petri nets which could be related in a clear and direct way. Intuitively, the difficulties encountered when matching temporal logics and Petri nets stem from the fact that the former are structured using composition operators, whereas the latter, in general, are not.

A notable exception is the Box Algebra (BA) [1] which supports Petri nets built using composition operators inspired by common programming constructs such as sequence, iteration, parallel composition and choice. Each Petri net box\((E)\) is derived from a box expression \(E\) using a compositional mapping box\((.)\). It is, therefore, natural to seek a temporal logic matching or supporting this particular set of composition operators. When looking at the existing temporal logics from this point of view, it was remarkable to realise that ITL (Interval Temporal Logic) is based on almost exactly the same set of programming constructs. As a result, we set out to explore the possibility of building a semantical bridge between temporal logics and Petri nets using two concrete formalisms, viz. ITL and BA. In this way, one should ultimately be able to take advantage of the individual strengths of these two formalisms, such as the analysis of systems with infinite data domains and fairness-related properties [16] for ITL, and unfolding based partial order model checking and invariant analysis for BA.

In concrete terms, our aim is to solve the following problem:

Given a BA expression \(E\) and a logic formula \(\psi\) expressed in ITL syntax, provide a translation \(\text{itl}(E)\) into ITL such that box\((E)\) satisfies \(\psi\) iff \(\text{itl}(E) \supset \psi'\) holds true, where \(\psi'\) is a suitably adjusted \(\psi\).

In this paper, we provide a syntax-driven translation \(\text{itl}(.)\) for the core BA [1] syntax comprising parallel composition, sequence, choice, synchronisation and iteration, but without considering data variables. Whereas translating the other control-flow operators is relatively straightforward, doing the same for synchronisation is much more involved and we rely here and adapt some ideas first formulated in [2] for the case of interprocess communication. We first consider a simple case of CCS-like binary synchronisation [9], and then extend this to a general multi-way synchronisation scheme [1]. The main result is the soundness of the proposed translation.

Throughout the paper \(\mathbb{N}\) denotes all positive integers, \(\mathbb{N}_0 = \mathbb{N} \cup \{0\}\) and \(\mathbb{N}_\omega = \mathbb{N}_0 \cup \{\omega\}\), where \(\omega\) denotes the first transfinite ordinal. We extend to \(\mathbb{N}_\omega\) the standard arithmetic comparison operators, assuming that \(\omega = \omega\) and \(n < \omega\), for all \(n \in \mathbb{N}_0\). Moreover, we define \(\leq\) as \(\leq \setminus \{\omega, \omega\}\). The concatenation operator for sequences will be denoted by \(\cdot\), and we will denote \(\omega^\omega = \{\omega, \omega, \ldots\}\) and \(\omega^+ = \{\varepsilon, \omega, \omega, \ldots\}\) (i.e., \(\omega^\omega\) comprises a single infinite sequence, and \(\omega^+\) an infinite number of finite sequences).
2 Box algebra with one-to-one communication

We first consider sba which is a simple sub-model of Box Algebra [1]. In particular, we will allow only one-to-one communication between concurrent sequential processes.

We assume a set of communication actions, each such action a having a unique conjugate action $\tilde{a}$ satisfying $\tilde{a} \neq a$ and $\tilde{\tilde{a}} = a$. Moreover, we allow synchronisation actions of the form $\tau_{(a,\tilde{a})}$ representing simultaneous execution of two conjugate communication actions, $a$ and $\tilde{a}$.

The syntax of sba expressions $E$ and sequential expressions $S$ is as follows:

$$
S ::= \text{stop} \mid a \mid S;S' \mid [S \circ S' \circ S''] \mid S \boxdot S'
$$

$$
E ::= (S_1 \parallel S_2 \parallel \ldots \parallel S_k) \circ A
$$

where $a$ is a communication action, and $A$ is a set of communication actions including conjugates. We assume that in an sba expression

$$
E = (S_1 \parallel S_2 \parallel \ldots \parallel S_k) \circ A
$$

we have $A_i \cap A_j = \emptyset$, for $i \neq j$, where each $A_i$ is the set of communication actions occurring in $S_i$.

In the above syntax, $\text{stop}$ stands for a blocked process, $S \circ S'$ for choice composition, $S;S'$ for sequential composition, $[S \circ S' \circ S'']$ for a loop (with an initial part $S$, iterated part $S'$, and terminal part $S''$), and finally $(S_1 \parallel S_2 \parallel \ldots \parallel S_k) \circ A$ for parallel composition of $k$ sequential processes. The scooping part, $\circ A$, of the expression $E$ in (1) synchronises conjugate communication actions in $A$ belonging to different sequential processes, leaving the non-synchronised actions (i.e., those not belonging to $A$) intact.

Figure 1 shows as sba expression $(S_1 \parallel S_2 \parallel S_3) \circ A$ modelling a system consisting of two one-place buffer processes, $S_1$ and $S_2$, and a merge process, $S_3$. Buffers $S_1$ and $S_2$ respectively use actions $a$ and $b$ to receive signals which are then send off to the merge process using the $c$ and $d$ actions. The merge process uses the conjugate actions, $\tilde{c}$ and $\tilde{d}$, to receive forwarded signals which are then passed on using the $e$ signal. The scooping part, $\circ A$, affects inter-process communication. (Note that the actions $c$, $\tilde{c}$, $d$ and $\tilde{d}$ cannot be executed individually.) The three processes are started up using the actions $\text{start}_i$ which are executed once.

Remark 1. The syntax of sba expression $E$ in (1) incorporates two specific restrictions:

- Only communication actions are used within sequential sub-expressions.

This is not a real problem as local (non-synchronised) actions are basically those communication actions which do not appear in $A$, and silent actions may be simulated by fresh communication actions.
Fig. 1. Two one-place buffers and merge processes.

- No action appears in more than one sequential expression $S_i$.
  If this does not hold, we can take any communication action $a \in A_i \cap A$ and then replace:
  - each occurrence of $a$ by $\text{stop} \square a_{i; j_i} \square \ldots \square a_{i; j_i}$ within $S_i$, and
  - $a$ by $a_{i_1; j_1}, \ldots, a_{i_k; j_k}$ within $A_i$.
  where $\{j_1, \ldots, j_k\} = \{j \mid j \neq i \land \hat{a} \in A_j\}$.

Box algebra semantics

The semantics of SBA expressions is given through a mapping into Petri nets called boxes.

A box is a tuple $\Sigma = (P, T, F, \ell, M_0)$ where $P$ and $T$ are disjoint finite sets of respectively places and transitions; $F \subseteq (P \times T) \cup (T \times P)$ is a flow relation; $\ell$ is a labelling function for places and transitions such that $\ell(p) \in \{e, i, x\}$, for every place $p \in P$, and $\ell(t)$ is an action, for every transition $t \in T$; and $M_0 \subseteq P$ is an (initial) marking. In general, any subset of $P$ is a marking.

We adopt the standard rules about representing nets as directed graphs. If the labelling of a place $p$ is $e$, $i$, or $x$, then $p$ is an entry, internal or exit place, respectively.

For every place (transition) $x$, we use $^x x$ to denote its pre-set, i.e., the set of all transitions (places) $y$ such that there is an arc from $y$ to $x$, that is, $(y, x) \in F$. The post-set $^x x$ is defined in a similar way. The pre- and post-set notation extends in the usual way to sets $R$ of places and transitions, e.g., $^x R = \bigcup_{r \in R} ^x r$. By convention, $^x \Sigma$ and $^x \Sigma$ denote respectively the sets of entry and exit places of $\Sigma$.

We now introduce operators on boxes corresponding to operators used in SBA expressions. Let $\Sigma_i$ (for $i = 1, 2, 3$) be boxes with disjoint sets of nodes satisfying $|^x \Sigma_i| = |^x \Sigma_i| = 1$ and $M_0 = \emptyset$. Then we have the following (below $p$ and $p'$ are fresh places):
\[ - \Sigma_1 \sqcup \Sigma_2 = (P, T, F, \ell, \emptyset) \text{ where:} \]
\[
P = P_1 \cup P_2 \setminus (\Sigma_1 \cup \Sigma_2 \cup \{p\}) \\
T = T_1 \cup T_2 \\
F = (F_1 \cup F_2)(P \times T) \cup \{p\} \\
\ell = (\ell_1 \cup \ell_2)_{p \mapsto T} \cup \{p \mapsto i\}.
\]
\[ - \Sigma_1 ; \Sigma_2 = (P, T, F, \ell, \emptyset) \text{ where:} \]
\[
P = P_1 \cup P_2 \setminus (\Sigma_1 \cup \Sigma_2 \cup \{p\}) \\
T = T_1 \cup T_2 \\
F = (F_1 \cup F_2)(P \times T) \cup \{p\} \times (\Sigma_2 \cup \Sigma_1) \times \{p\} \\
\ell = (\ell_1 \cup \ell_2)_{p \mapsto T} \cup \{p \mapsto i\}.
\]
\[ - [\Sigma_1 \circ \Sigma_2 \circ S_3] = (P, T, F, \ell, \emptyset) \text{ where:} \]
\[
P = P_1 \cup P_2 \cup P_3 \setminus (\Sigma_1 \cup \Sigma_2 \cup \Sigma_3 \cup \{p\}) \\
T = T_1 \cup T_2 \cup T_3 \\
F = (F_1 \cup F_2 \cup F_3)(P \times T) \cup \{p\} \times (\Sigma_2 \cup \Sigma_3) \times \{p\} \\
\ell = (\ell_1 \cup \ell_2 \cup \ell_3)_{p \mapsto T} \cup \{p \mapsto i\}.
\]

Moreover, for any box \( \Sigma \), \( \overline{\Sigma} \) is \( \Sigma \) with the initial marking set to \( \Sigma \).

The semantics of a box \( \Sigma \) is given through its step sequences. A set of transitions \( U \), called a step, is enabled at a marking \( M \) if \( \Sigma U \subseteq M \) and \( \Sigma U \cap \Sigma t = \emptyset \) for all distinct \( t, u \in U \). An enabled step \( U \) can be executed leading to a marking \( M' \) given by \( M' = M \setminus U \). We denote this by \( M[\Sigma U] \).

As far as a box \( \Sigma \) is concerned, only step sequences which start from its default initial marking \( \Sigma U \) need to be considered. We will assume that each such step sequences is infinite which is a harmless requirement as any finite step sequence can be extended by an infinite sequence of empty steps (note that \( M[\emptyset] M \) for every marking \( M \)). In addition, we will single out a set of finite step sequences which lead from the default initial marking \( \Sigma U \) to the default final marking \( \Sigma \). Intuitively, each such step sequence corresponds to a terminated execution of the box.

A step sequence of a box \( \Sigma \) is any infinite sequence of steps \( \gamma = U_1 U_2 \ldots \) such that there are markings \( M_1, M_2 \ldots \) satisfying

\[ \Sigma[U_1]M_1[U_2]M_2 \ldots \]

We denote this by \( \gamma \in \text{step}(\Sigma) \). Moreover, a terminated step sequence of \( \Sigma \) is a finite sequence of steps \( \gamma = U_1 \ldots U_m \) such that there are markings \( M_1, \ldots, M_{m-1} \) satisfying

\[ \Sigma[U_1]M_1[U_2] \ldots M_{m-1}[U_m] \Sigma \]


We denote this by $\gamma \in \text{lstep}(\Sigma)$.

Step sequences built of sets of transitions are low-level descriptions of executed behaviours. A more abstract (and practically relevant) view is provided by step sequences built of steps of labels of executed transitions. Hence, for any finite or infinite step sequence $\gamma$ as defined above, we will use $\ell(\gamma)$ to denote a sequence of multisets of labels obtained from $\gamma$ by replacing each step $U_i$ by the multiset $\Gamma_i$ of labels of the transitions belonging to $U$. We then define:

$$\text{lstep}(\Sigma) = \ell(\text{step}(\Sigma)) \quad \text{and} \quad \text{ulstep}(\Sigma) = \ell(\text{ulstep}(\Sigma)).$$

**From box expressions to boxes**

We define a mapping box from SBA expressions to boxes compositionally. First, for the blocked expression stop and any communication action $a$:

$$\text{box}(a) = \{(p, p'), \{a, t\}, \{p \mapsto e, t_a \mapsto a, p' \mapsto x\}, \emptyset\}$$

$$\text{box}(\text{stop}) = \{(p, p'), \emptyset, \emptyset, \{p \mapsto e, p' \mapsto x\}, \emptyset\}$$

and, for any sequential expressions $S_1$, $S_2$ and $S_3$:

$$\text{box}(S_1 ; S_2) = \text{box}(S_1) ; \text{box}(S_2)$$

$$\text{box}(S_1 \sqcup S_2) = \text{box}(S_1) \sqcup \text{box}(S_2)$$

$$\text{box}([S_1 \sqcup S_2 \sqcup S_3]) = \text{box}(S_1) \sqcup \text{box}(S_2) \sqcup \text{box}(S_3)$$

Then, for an SBA expression $E$ as in (1), the net $\text{box}(E)$ is obtained by:

- creating $\Sigma$ which is a disjoint union of $\text{box}(S_i)$, for $i = 1, \ldots, k$.
- creating $\Sigma'$ from $\Sigma$ by adding fresh transitions $t = t_{(u, v)}$, where $u$ is a transition in $\text{box}(S_i)$ and $v$ is a transition in $\text{box}(S_j)$ ($i \neq j$), $t_i(u) = a \in A$ and $\ell_i(u) = \hat{a}$. The label of $t$ is $\tau(a, \hat{a})$, and it inherits the connectivity of $u$ and $v$, i.e., $*t = *u \cup *v$ and $*t = u^* \cup v^*$.
- deleting all transition labelled by the communication actions in $A$.

We also define, for any SBA expression $F$:

$$\text{lstep}(F) = \text{lstep}(\text{box}(F)) \quad \text{and} \quad \text{ulstep}(F) = \text{lstep}(\text{box}(F)).$$

Behaviours of boxes obtained through the above translation from SBA expressions exhibit clear compositional properties [1]. Figure 2 provides a full characterisation of the behaviours of sequential SBA expressions whereas for concurrent SBA expressions we have the following.

**Proposition 1.** Let $E$ be an SBA expression as in (1). Then:

- a finite sequence $\gamma = \Gamma_1 \ldots \Gamma_m$ belongs to $\text{lstep}(E)$ iff for $j = 1, \ldots, k$:

  $$\Gamma_j^i \ldots \Gamma_j^i \in \text{lstep}(S_j)$$

  where $\Gamma_j^i = \Gamma_j |_{\Gamma_j^i} + \{a \in A_j \mid \tau(a, \hat{a}) \in \Gamma_j^i\}$, for $i = 1, \ldots, m$. 

- an infinite sequence $\gamma = \Gamma_1 \Gamma_2 \ldots$ belongs to $\text{lstep}(E)$ iff for $j = 1, \ldots, k$:

$$\Gamma_1 \Gamma_2 \ldots \in \text{lstep}(S_j)$$

where $\Gamma_i = \Gamma_i|_{A_j} + \{ a \in A_j \mid \tau_{(a, \bar{a})} \in \Gamma_1 \}$, for $i = 1, 2, \ldots$.

Note: $+$ and $|_{A_j}$ respectively denote multiset sum and restriction. 

Figure 3 illustrates the construction of a box representing an sfa expression of Figure 1.

3 Interval Temporal Logic

We now provide the syntax and semantics of a small fragment of ITL. The chosen fragment includes only those constructs (basic and derived) which are used in the subsequent translation of sfa expressions.

The formulas of the fragment of the ITL logic we need are defined below, where $V$ is a set of boolean variables, $v \in V$ and $V' \subseteq V$:

$$\phi ::= \text{flip}(v) \mid \text{keep}(V') \mid \phi \land \phi' \mid \phi \lor \phi' \mid \phi \hat{\tau} \phi' \mid \phi^* \mid \text{inf}$$

Note that $\text{flip}(v)$ inverts the value of boolean variable $v$ over a unit interval, $\text{keep}(V')$ keeps the value of variables in $V'$ over a unit interval, $\hat{\tau}$ is a weak version of the standard sequential composition operator (denoted by $;$ and called chop), and $\text{inf}$ indicates an infinite interval.

A state is a mapping which assigns values to the (boolean) variables $V$, and an interval $\sigma$ is a possibly infinite non-empty sequence of states. Its length, $|\sigma|$, is $\omega$ if $\sigma$ is infinite, and otherwise its number of states minus 1. To simplify

$$\text{lstep}(\text{stop}) = \emptyset$$

$$\text{lstep}(a) = \emptyset^* \circ \{(a)\} \circ \emptyset^*$$

$$\text{lstep}(S_1 ; S_2) = \text{lstep}(S_1) \circ \text{lstep}(S_2)$$

$$\text{lstep}(S_1 \cap S_2) = \text{lstep}(S_1) \cup \text{lstep}(S_2)$$

$$\text{lstep}([S_1 \oplus S_2 \oplus S_3]) = \text{lstep}(S_1) \circ \text{lstep}(S_2)^* \circ \text{lstep}(S_3)$$

$$\text{lstep}(\text{stop}) = \emptyset^\omega$$

$$\text{lstep}(a) = \emptyset^\omega \circ \{(a)\} \circ \emptyset^\omega$$

$$\text{lstep}(S_1 ; S_2) = \text{lstep}(S_1) \cup \text{lstep}(S_1) \circ \text{lstep}(S_2)$$

$$\text{lstep}(S_1 \cap S_2) = \text{lstep}(S_1) \cup \text{lstep}(S_2)$$

$$\text{lstep}([S_1 \oplus S_2 \oplus S_3]) = \text{lstep}(S_1) \cup \text{lstep}(S_2)^* \circ (\text{lstep}(S_2) \cup \text{lstep}(S_3))$$

Fig. 2. Properties of behaviours of sequential sfa expressions.
definitions, we will denote $\sigma$ as $\langle \sigma_0, \sigma_1, \ldots, \sigma_{|\sigma|} \rangle$, where $\sigma_{|\sigma|}$ is undefined if $\sigma$ is infinite. With such a notation, for $0 \leq i \leq j \leq |\sigma|$: $\sigma_{i..j} = \langle \sigma_i, \ldots, \sigma_j \rangle$ and $\sigma^i = \langle \sigma_0, \ldots, \sigma_i \rangle$ and $\sigma^{(i)} = \langle \sigma_i, \ldots, \sigma_{|\sigma|} \rangle$

The meaning of formulas is given by the satisfaction relation defined as follows:

- $\sigma \models \text{flip}(v)$ iff $|\sigma| = 1$ and $\sigma_1(v) = \neg \sigma_0(v)$.
- $\sigma \models \text{keep}(\{v_1, \ldots, v_m\})$ iff $|\sigma| = 1$ and $\sigma_1(v_i) = \sigma_0(v_i)$, for $i = 1, \ldots, m$.
- $\sigma \models \phi \lor \phi'$ iff $\sigma \models \phi$ or $\sigma \models \phi'$.
- $\sigma \models \phi \land \phi'$ iff $\sigma \models \phi$ and $\sigma \models \phi'$.
- $\sigma \models \phi \triangleright \phi'$ iff one of the following holds:
  - $|\sigma| = \omega$ and $\sigma \models \phi$.
  - there is $r \leq |\sigma|$ and $\sigma^r \models \phi$ and $\sigma^{(r)} \models \phi'$.
- $\sigma \models \phi^*$ iff one of the following holds:
  - $|\sigma| = 0$.
  - there are $0 = r_0 \leq r_1 \leq \ldots \leq r_{n-1} \leq r_n = |\sigma|$ such that, for all $1 \leq l \leq n$, $\sigma_{r_{l-1..r_l}} \models \phi$. 

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{figure3}
\caption{Top: unsynchronised unmarked boxes for $S_1, S_2$ and $S_3$; bottom: complete box semantics of the three-process system with the default initial marking. All the $i$ labels of internal places are omitted.}
\end{figure}
\[ |\sigma| = \omega \text{ and there are infinitely many integers } 0 = r_0 \leq r_1 \leq \ldots \text{ such that } \lim_{i \to \infty} r_i = \omega \text{ and for all } l \geq 1, \sigma_{r_{l-1} \ldots r_l} = \phi. \]

- \( \sigma \models \inf \iff |\sigma| = \omega. \)

Note that \( \phi \models \phi' \) is equivalent to the formula \( \phi \land \inf \lor \phi \land \phi' \).

### 4 From simple box algebra to IL

We now present a translation from an SBA expression \( E \) defined as in (1) into semantically equivalent IL formula.

A key idea inspired by [2] is to represent each of the actions \( a \) appearing in \( E \) by a separate boolean variable \( v_a \), and then to model an execution of \( a \) by the change of the value of \( v_a \). In the case of synchronisation between \( a \) and \( \bar{a} \), the two corresponding variables, \( v_a \) and \( v_{\bar{a}} \), have to change their values simultaneously.

In what follows, the set of variables corresponding to each \( A_i \) (actions occurring within \( S_i \)) are denoted by \( V_i \), and no other variables are used. The translation is then given by:

\[
\text{itl}(E) = \text{itl}(S_1) \land \ldots \land \text{itl}(S_k)
\]

where we have the following (below \( i = 1, \ldots, k \), \( a \in A_i \cap A \) and \( b \in A_i \setminus A \)):

\[
\begin{align*}
\text{itl}_i(\text{stop}) &= \text{keep}(V_i)^* \land \inf \\
\text{itl}_i(a) &= \text{keep}(V_i)^* \land (\text{keep}(V_i \setminus \{v_a\}) \land \text{flip}(v_a) \land \text{flip}(v_{\bar{a}}))^* \\
\text{itl}_i(b) &= \text{keep}(V_i)^* \land (\text{keep}(V_i \setminus \{v_b\}) \land \text{flip}(v_b))^* \land \text{keep}(V_i)^* \\
\text{itl}_i(S; S') &= \text{itl}_i(S) \lor \text{itl}_i(S') \\
\text{itl}_i(S \circ S') &= \text{itl}_i(S) \land \text{itl}_i(S') \\
\text{itl}_i([S \circ S' \circ S'']) &= \text{itl}_i(S) \land \text{itl}_i(S') \land \text{itl}_i(S'').
\end{align*}
\]

Intuitively, the value of a variable \( v_a \) is kept unchanged, unless we simulate an execution of action \( a \) which results in flipping the value of \( v_a \). Moreover, synchronisation involving \( a \) and \( \bar{a} \) flips the values of both \( v_a \) and \( v_{\bar{a}} \). In this way, executions carried out in concurrent components are synchronised.

With each interval \( \sigma \) satisfying \( \sigma \models \text{itl}(E) \), we associate a sequence of multisets \( \gamma_\sigma = \Gamma_1 \ldots \Gamma_{|\sigma|} \), where each \( \Gamma_i \) is defined as follows:

\[
\Gamma_j = \{ a \mid a \notin A \land \sigma_{j-1}(v_a) \neq \sigma_j(v_a) \} + \{ \tau(a, a) \mid a \in A \land \sigma_{j-1}(v_a) \neq \sigma_j(v_a) \}.
\]

Similarly, with each interval \( \sigma \) satisfying \( \sigma \models \text{itl}(S_i) \), we associate a sequence of multisets \( \gamma_\sigma = \Gamma_1 \ldots \Gamma_{|\sigma|} \), where each \( \Gamma_j \) is defined as follows:

\[
\Gamma_j = \{ a \in A_i \mid \sigma_{j-1}(v_a) \neq \sigma_j(v_a) \}.
\]
We then define, for any \( \Pi I \) formula \( \phi \) appearing on the r.h.s. of the above translation:

\[
ltstep(\phi) = \{ \gamma \sigma \mid \sigma \models \phi \land |\sigma| < \omega \} \quad \text{and} \quad lstep(\phi) = \{ \gamma \sigma \mid \sigma \models \phi \land |\sigma| = \omega \}.
\]

Note that \( ltstep(\phi) \circ \varnothing^\omega \subseteq lstep(\phi) \).

**Proposition 2.** Let \( S \) be a sequential \( SBA \) expression. Then:

\[
lstep(itl(S)) = lstep(S) \quad \text{and} \quad lstep(itl(S)) = lstep(S).
\]

**Proof.** The proof proceeds by induction on the structure of \( S \), using Figure 2 and the properties of the logic operators.

Case 1: \( S = \text{stop} \). Then:

\[
ltstep(itl(\text{stop})) = \varnothing = lstep(\text{stop})
\]

\[
lstep(itl(\text{stop})) = \varnothing^\omega = lstep(\text{stop}).
\]

Case 2: \( S = a \). Then:

\[
ltstep(itl(a)) = \varnothing^* \circ \{\{a\}\} \circ \varnothing^* = lstep(a)
\]

\[
lstep(itl(a)) = \varnothing^\omega \cup \varnothing^* \circ \{\{a\}\} \circ \varnothing^\omega = lstep(a).
\]

Case 3: \( S = S_1; S_2 \). Then:

\[
lstep(itl(S_1; S_2)) = ltstep(itl(S_1)) \circ ltstep(itl(S_2))
\]

\[
= lstep(S_1) \circ lstep(S_2)
\]

\[
= lstep(S_1; S_2)
\]

\[
lstep(itl(S_1; S_2)) = lstep(itl(S_1)) \cup lstep(itl(S_1)) \circ lstep(itl(S_2))
\]

\[
= lstep(S_1) \cup lstep(S_1) \circ lstep(S_2)
\]

\[
= lstep(S_1; S_2).
\]

Case 4: \( S = S_1 \circ S_2 \). Then:

\[
lstep(itl(S_1 \circ S_2)) = ltstep(itl(S_1)) \cup ltstep(itl(S_2))
\]

\[
= lstep(S_1) \cup lstep(S_2)
\]

\[
lstep(itl(S_1 \circ S_2)) = lstep(itl(S_1)) \cup lstep(itl(S_2))
\]

\[
= lstep(S_1) \cup lstep(S_2)
\]

\[
= lstep(S_1 \circ S_2).
\]
Case 5: $S = [S_1 \oplus S_2 \oplus S_3]$. Then:

$$lstep(ill([S_1 \oplus S_2 \oplus S_3])) = lstep(ill(S_1)) \circ lstep(ill(S_2))^* \circ lstep(ill(S_3))$$
$$= lstep(S_1) \circ lstep(S_2)^* \circ lstep(S_3)$$
$$= lstep([S_1 \oplus S_2 \oplus S_3])$$

$$lstep(ill([S_1 \oplus S_2 \oplus S_3])) = lstep(ill(S_1)) \cup lstep(ill(S_2))^* \circ (lstep(ill(S_2)) \cup lstep(ill(S_3)))$$
$$= lstep(S_1) \cup lstep(S_2)^* \circ (lstep(S_2) \cup lstep(S_3))$$
$$= lstep([S_1 \oplus S_2 \oplus S_3]) .$$

This completes the proof. \[
\]

**Theorem 1.** Let $E$ be an sba expression as in (1). Then:

$$lstep(ill(E)) = lstep(E) \quad \text{and} \quad lstep(ill(E)) = lstep(E) .$$

**Proof.** Follows from Propositions 1 and 2, and the basic properties of the conjunction operator in ITL. Note that the assumption that no communication action occurs in more than one sequential expression $S_i$ is crucial to demonstrate the result. \[
\]

What we have just presented is one of possible ways of translating sba expressions into equivalent ITL formulas. Another possibility would be to enforce synchronisation between variables corresponding to conjugate communications globally using the $\boxdot$ operator of ITL:

$$ill(E) = ill_1(S_1) \land \ldots \land ill_k(S_k) \land \bigwedge_{a \in A} \Box v_a = v_{\bar{a}}$$

and then to simplify the rest of the translation (below $b \in A_i$):

$$ill_i(\text{stop}) = \text{keep}(V_i)^* \land \text{inf}$$
$$ill_i(b) = \text{keep}(V_i)^* \land \text{keep}(V_i \setminus \{v_b\} \land \text{flip}(v_b))^* \land \text{keep}(V_i)^*$$
$$ill_i(S \cdot S') = ill_i(S) \land ill_i(S')$$
$$ill_i(S \boxdot S') = ill_i(S) \lor ill_i(S')$$
$$ill_i([S \oplus S' \oplus S'']) = ill_i(S) \land ill_i(S') \land ill_i(S'')$$
Yet another possibility would be to use a single variable $v_{(a,\hat{a})}$ to represent both $a$ and $\hat{a}$, leading to the following translation (below $a \in A_1 \cap A$ and $b \in A_1 \setminus A$):

$$
\begin{align*}
\text{ilt}(E) &= \text{ilt}_1(S_1) \land \ldots \land \text{ilt}_k(S_k) \\
\text{ilt}_i(\text{stop}) &= \text{keep}(V_i)^* \land \text{inf} \\
\text{ilt}_i(a) &= \text{keep}(V_i)^* \top \\
&\quad (\text{keep}(V_i \setminus \{v_{(a,\hat{a})}\}) \land \text{flip}(v_{(a,\hat{a})})) \top \text{keep}(V_i)^* \\
\text{ilt}_i(b) &= \text{keep}(V_i)^* \top (\text{keep}(V_i \setminus \{v_b\}) \land \text{flip}(v_b)) \top \text{keep}(V_i)^* \\
\text{ilt}_i(S; S') &= \text{ilt}_i(S) \oplus \text{ilt}_i(S') \\
\text{ilt}_i(S \sqcup S') &= \text{ilt}_i(S) \lor \text{ilt}_i(S') \\
\text{ilt}_i([S \otimes S' \otimes S'']) &= \text{ilt}_i(S) \ominus \text{ilt}_i(S') \ominus \text{ilt}_i(S'')
\end{align*}
$$

The latter alternative will be adopted in the translation of more complicated BA expressions described in the next section.

## 5 Box algebra with general synchronisation

We now consider BA equipped with a powerful operator of general synchronisation which subsumes a majority of those usually employed by process algebras.

Let $L$ be a set of actions. A synchronisation relation is $\rho \subseteq L^{+} \times L$ such that if $(a_1, \ldots, a_n, a) \in \rho$ then $a_1, \ldots, a_n$ represents actions which can be synchronised to yield a new action $a$. The general syntax of BA expressions $E$ and sequential BA expressions $S$ is as follows:

$$
\begin{align*}
S &::= \text{stop} \mid a \mid S; S' \mid [S \otimes S' \otimes S''] \mid S \sqcup S' \\
E &::= (S_1 \parallel S_2 \parallel \ldots \parallel S_k)[\rho]
\end{align*}
$$

where $a \in L$ is an action, and $\rho$ is a synchronisation relation. There is no restriction on the presence of actions in different sequential subexpressions of $E$ as in the case of SBA.

The translation for a sequential expression $S$ is the same as in the simpler case. Moreover, given a BA expression

$$
E = (S_1 \parallel S_2 \parallel \ldots \parallel S_k)[\rho]
$$

we proceed as follows:

- For every $\xi = (a_1, \ldots, a_n, a) \in \rho$ we denote:

$$
\Delta_\xi = \{k_1 \ldots k_n \mid \forall i \leq n : a_i \in A_{k_i} \land \forall i \neq j : k_i \neq k_j\}.
$$

Intuitively, $\Delta_\xi$ denotes all combinations of sequential processes which can potentially generate synchronisations captured by $\xi$.

- $\text{box}(E)$ is obtained by:
  * creating $\Sigma$ which is a disjoint union of $\text{box}(S_i)$, for $i = 1, \ldots, k$. 


( [ start ⊕ (a;c) ⊕ stop ] || with ρ = { (start, start, start), start),
[ start ⊕ (f;c) ⊕ stop ] || (f, f'), (f, g),
[ start ⊕ (b;c') ⊕ stop ] ) [ρ] (a, a), (b, b), (c, e) }

Fig. 4. A system with many-to-one communication.

• creating Σ' from Σ by adding a transition t = t₁...ₙ whenever:

\[ ξ = (a₁, ..., aₙ, a) \in ρ \text{ and } k₁...ₙ \in Δξ \]

and, for all i ≤ n, tᵢ belongs to box(Sᵢ) and has the label aᵢ. The label of t is a, \( t^* = t₁ \cup ... \cup tₙ \) and \( t^* = t₁^* \cup ... \cup tₙ^* \).

• deleting all transitions coming from Σ.

The translation from the RA expressions to boxes follows the same pattern as in the case of SBA expressions, and its result is illustrated in Figure 4.

From box algebra to ITL

In the translation from RA expressions as in (2) to ITL, for every synchronisation pattern \( ξ = (a₁, ..., aₙ, a) \in ρ \), we will use distinct variables

\[ V_ξ = \{ v_{k₁...ₙ}^ξ | k₁...ₙ \in Δξ \} \]

corresponding to different combinations of sequential sub-expressions which may realise synchronisations captured by ξ. Moreover, for every \( i \leq k \) and \( b \in A_i \),

\[ V_i^b = \{ v_{k₁...ₙ}^{(a₁...ₙ,a)} | \exists j : k_j = i \wedge a_j = b \} \]

are all the variables corresponding to potential synchronisations in which b occurring in \( S_i \) can participate. The translation of RA expressions is then defined.
in the following way (below $i = 1, \ldots, k$, $b \in A_i$ and $V_i = \bigcup_{b \in A_i} V_i'$):

\[
\begin{align*}
  \text{itt}(E) & = \text{itt}_1(S_1) \land \ldots \land \text{itt}_k(S_k) \\
  \text{itt}_i(\text{stop}) & = \text{keep}(V_i)^* \land \text{inf} \\
  \text{itt}_i(b) & = \text{keep}(V_i)^* \land \left( \bigvee_{v \in V_i} \text{keep}(V_i \setminus \{v\}) \land \text{flip}(v) \right) \land \text{keep}(V_i)^* \\
  \text{itt}_i(S_1 \cdot S') & = \text{itt}_i(S_1) \land \text{itt}_i(S') \\
  \text{itt}_i(S \parallel S') & = \text{itt}_i(S) \lor \text{itt}_i(S') \\
  \text{itt}_i([S \otimes S' \otimes S'']) & = \text{itt}_i(S) \land \text{itt}_i(S') \land \text{itt}_i(S'')
\end{align*}
\]

Note that a synchronisation can now involve more than two sequential sub-components and is realised by a single variable which may appear in several different sub-formulas. The interpretation of satisfying intervals in terms of step sequences of multisets of actions requires only a small modification of the one used previously.

Properties of the newly defined translation are very much the same as in the simpler case. Crucially, we have the following.

**Theorem 2.** Let $E$ be a $\lba$ expression as in (2). Then:

\[
\text{ltstep}(\text{itt}(E)) = \text{ltstep}(E) \text{ and } \text{ltstep}(\text{itt}(E)) = \text{ltstep}(E).
\]

**Proof.** Similar as in the case of $\sba$ expression. \qed

6 Conclusions

In the past, various kinds of logics have been used as formalism for expressing correctness properties of systems specified using Petri nets. When it comes to the relationship between logics and Petri net, we feel that the work on the connections between linear logic [5] and Place Transition nets was the closest one. However, the main concern there was the handling of multiple token occurrences in net places whereas boxes are safe nets. Another way in which logics and Petri nets were discussed was reported in [14] which provided a characterisation of Petri net languages in terms of second-order logical formulas.

The results presented in this paper demonstrate that one can develop a very close structural connection between $\lba$ and $\ltl$. It is therefore important to further investigate the extent to which such a connection could be generalised and exploited. In particular, we plan to investigate what is the subset of $\ltl$ which can be modelled by $\lba$. A longer term goal is the development of a hybrid verification methodology combining $\ltl$ and $\lba$ techniques. For example, sequential algorithms and data structures could be treated by $\ltl$ techniques [11], while intensive parallel or communicating aspects of systems could be treated by net unfoldings [4, 6] or other Petri net techniques [15].
Acknowledgement

This research is supported by the 973 Program Grant 2010CB328102, NSFC Grants 60910004 and 60873018, and EPSRC VERDAD project.

References

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