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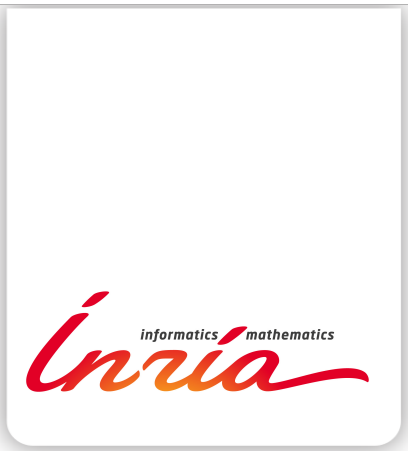

Using counters for absence prediction in Esterel

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Abstract: **Esterel** is a **synchronous programming language** historically defined for system control, well suited to react in parallel to external sensors, intensively used in avionics. Recently, with the incoming of the orchestration language HipHop, a domain-specific language of the multi-tier language **Hop**, Esterel is used to manage Web requests. In this context, where orchestration programs are dynamically generated, long compilation preamble to computation must be avoided and a simple and fast interpreter is preferred. This paper presents such an interpreter. Esterel's processes communicates through signals and one particularity of this language is its ability to instantaneously react to the absence of a signal. In this paper we present a static analysis which allows the interpreter to predict the absence of a signal.

Key-words: synchronous languages, instantaneous reaction to absence, interpretation, continuations, formal proof

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Utilisation de compteurs pour la prédiction de l'absence en Esterel

Résumé : Esterel est un langage synchrone historiquement défini pour les systèmes de contrôle, particulièrement adapté pour réagir en parallèle à des évènements externes, utilisé intensivement dans l'avionique. Récemment, avec l'arrivée du langage d'orchestration HipHop, un sous-langage dédié de Hop, l'approche Esterel est utilisée pour synchroniser des requêtes Web. Dans ce contexte, où les programmes d'orchestration sont générés dynamiquement, les longs préambules de compilation sont à éviter et l'utilisation d'interprètes simples et rapides devient souhaitable. Cet article présente un tel interprète. Les processus Esterel communiquent au travers de signaux et l'une des particularités de ce langage est sa capacité de réagir instantanément à l'absence d'un signal. Dans cet article, nous présentons une analyse statique qui permet à l'interprète de prédire l'absence d'un signal.

Mots-clés : langages synchrones, réaction immédiate à l'absence, interprétation, continuation, preuve formelle

1 Introduction

We informally present the Esterel programming language focusing on the notion of process. An Esterel program is made of several processes, each of them executing statements. The processes communicate via *signals*. A process can raise a signal S with the statement "**emit** S". Once emitted, a signal is visible by all the processes. Signals can be tested with the statement "**present** S **then** P_1 **else** P_2 **end**". If the signal S is already emitted, the evaluation continues with P_1. We will detail this statement latter in this section. A process can emit a signal but cannot reset it, as if it was never been emitted. The only way to reset a signal is to erase all signals at the same time defining a notion of *instant*. All processes must cooperate to reach the end of one instant. Individually, a process may decide to finish its instant by executing the statement "**pause**", and it is when all processes have executed a pause or have finished their execution that the current instant is globally closed.

Processes are generated via a statement "$P_1 \| P_2$". At runtime two processes are created, one executing P_1, the other executing P_2, the main process waiting for the completion of its two sub-processes before continuing its own execution. The execution of P_1 and P_2 may take several instants. The two sub-processes must join on completion before resuming the execution of their creator.

Coming back to the "**present**" statement, if one process have to evaluate a statement "**present** S **then** P_1 **else** P_2 **end**", the signal S may be in an undefined state : not yet emitted but a concurrent process having the opportunity to emit it. In this case, the current process will wait until the status of the signal is defined. A signal is present once an "**emit** " is evaluated. Conversely, a signal is absent if and only if it is not emitted in the instant. So, to predict the absence of a signal in a middle of an instant, we have to look in the future to prove that no process will emit it. In this work, to detect the absence of a signal, we maintain a prediction of how many times a signal may be emitted by all processes until the end of the current instant. The prediction is correct if, when the count of a signal reaches zero, then the signal cannot be emitted during the instant.

The prediction is over-estimated. If for a signal the prediction gives a value of n, this means that, for the rest of the instant, the signal can be emitted *at most* n times. For a "**present**" statement, both the "**then**"and the "**else**" part of the statement are considered as reachable when predicting. Dynamically, when one branch of control will be taken, the potentially emitted signals of the other branch is erased from the prediction.

Let's consider a simple example ("P_1;P_2" is the sequence of the two statements P_1 and P_2) :

```
   (P; emit S)
|| present S then emit Y else emit N end
```

If the statement P is a "**pause**", then "**emit** S" will be done in the next instant and the count associated to S is zero; thus, the signal can be considered as absent and the signal N is emitted. If the statement P is "**nothing**" (a statement that does nothing, runs in no time without emitting any signal), the count associated to S is one and the signals S and Y will be emitted. Finally, consider the case where the statement P is "**present** I **then pause else nothing end**"; statically we cannot know if "**emit** S" will be executed in the current instant or in the next instant; in this case the prediction for S is one, but dynamically, if the signal I becomes present, the count for S is decremented by one, reaching the value 0, and the signal S can safely be considered as absent.

In this paper we define a static analysis that creates these predictions and attaches them to programs. We also describe how an interpreter uses this information to predict the absence of signals. Finally, we prove that signals predicted absent are not emitted during the instant.

The rest of the paper is organised as follows. Section 2 finishes to introduce informally the semantics of Esterel. Section 3 defines the static analysis which associates to the nodes of the abstract syntax the prediction of how many times each signal can be emitted. Section 4 defines the interpretation of the decorated nodes and how the predictions are used and dynamically updated. In this section, we also discuss on continuations which implies trampolines techno-

nothing	Nop	Nop
pause	$Pause$	$Pause^+(\sigma)$
emit S	$Emit(S)$	$Emit(S)$
P_1;P_2	$Seq(P_1, P_2)$	$Seq(P_1, P_2)$
$P_1 \| P_2$	$Par(P_1, P_2)$	$Par^+(P_1, P_2, \sigma)$
present S **then** P_1 **else** P_2 **end**	$If(S, P_1, P_2)$	$If^+(S, P_1, \sigma_1, P_2, \sigma_2)$
loop P **end**	$Loop(P)$	$Loop^+(P, \sigma)$
trap T **in** P **end**	$Trap(T, P)$	$Trap^+(T, P, \sigma)$
exit T	$Exit(T)$	$Exit(T)$
signal S **in** P **end**	$Signal(S, P)$	$Signal^+(S, P, n)$

Figure 1: Syntax, AST and decorated AST of core Esterel

logy, show the synchronisation reclaimed by the parallelism operator, and present some implementations of the waiting processes. In Section 5 we formally describe the analysis and the evaluator expressed in the **Coq**[1] system. With this formal specification, we prove the correctness of the analysis : if a signal is considered absent by the evaluator at one instant, then this signal will be not emitted during this instant. In Section 6 we briefly discuss related work before concluding.

2 Syntax and informal semantics

The syntax of core Esterel is described in the first column of Figure 1. The six first statements were presented in the introduction. Esterel has only one basic notion of recursion which is the infinite loop : "**loop** P **end**" is the infinite repetition of the statement P. The statement "**trap** T **in** P **end**" allows the control to exit from a loop, it defines an escape point named T during the execution of P. The execution of P is aborted with a statement "**exit** T". Evaluating an "**exit** T" cannot run directly to its binder and has to join on each parallel computation for possible deeper exits. Let's take the example :

```
trap T in
    trap U in
        (... exit U ...)
    || Q
end end
```

When the left process reaches the "`exit U`", it has to wait the completion of `Q`. The right process may evaluate an "`exit T`" and thus overtake the *pending* "`exit U`".

For example, the couple "**trap in end**"/"**exit** " can be used to wait until a signal S is emitted :

```
trap T in
    loop
        present S then exit T else pause
end end end
```

The statement "**signal** S **in** P **end**" defines a local signal named S during the execution of P. In the first part of the paper we will assume a full renaming of local signals. A program like :

```
    emit S
 || signal S in P end
 || signal S in Q end
```

will be rewritten as :

```
    emit S
 || signal S1 in P1 end
 || signal S2 in Q2 end
```

with `P1` (resp. `Q2`) being the statement `P` (resp. `Q`) where all free occurrences of `S` are replaced by `S1` (resp. `S2`). Note that, even renamed with a fresh signal name, a local signal declaration cannot be shifted

[1] In this paper, all underlined bold **words** can be found in Wikipedia.

up to the top of the program due to the *schizophrenic* pattern, like in the following program :

```
loop
    signal S in
        present S
           then emit Y
           else emit N end;
        pause;
        emit S
end end
```

The local signal definition of `S` prevents to see the signal as present. Without the local definition, the signal, except for the first instant, will be always present. We will come back to the *schizophrenic* pattern in the next section.

3 Emit prediction

We will define a static analysis that associates to each statement the signals that can be emitted during the evaluation of this statement until the next instant, i.e. until reaching a pause or the end of the computation. As a simple example, for the statement `emit A; emit B; emit A; pause; ...`, we will associate the fact that `A` will be emitted 2 times and `B` once. These values are named *predictions* and noted $\sigma = \{A \to 2, B \to 1\}$, where σ ia a prediction function from signal names to integers. If $\bullet$ is a function on integers, we note $\sigma_1 \bullet \sigma_2$ the function $\lambda s.\sigma_1(s) \bullet \sigma_2(s)$[2].

In the previous example, the prediction is exact, but there are cases where we cannot compute exact predictions. For example in the statement "`present S then emit A else emit B end`", we cannot predict statically if either `A` or `B` will be emitted. We thus compute an over-estimation by considering that both `A` and `B` can be emitted. In other words, for a `present` statement, the prediction is the sum of the predictions of its two branches. At runtime, when the state of the signal `S` is clearly present or absent, the prediction of the branch that is not evaluated has to be decremented.

[2]We use the **λ-calculus** notation for anonymous function : $\lambda x.B$ is the function with a formal parameter x and computing B when called.

$analyze(P, \tau, \sigma) \stackrel{\triangle}{=}$ **match** P **with**
10 $Nop \Rightarrow \sigma, Nop$
20 $Pause \Rightarrow \emptyset, Pause^+(\sigma)$
30 $Emit(S) \Rightarrow\ (\sigma + \{S \to 1\}), Emit(S)$
40 $Seq(P_1, P_2) \Rightarrow \sigma_1, Seq(P_1^+, P_2^+)$
41 **where** $\sigma_1, P_1^+ = analyze(P_1, \tau, \sigma_2)$
42 **where** $\sigma_2, P_2^+ = analyze(P_2, \tau, \sigma)$
50 $Par(P_1, P_2) \Rightarrow (\sigma_1 + \sigma_2), Par^+(P_1^+, P_2^+, \sigma)$
51 **where** $\sigma_1, P_1^+ = analyze(P_1, \tau, \sigma)$
52 **and** $\sigma_2, P_2^+ = analyze(P_2, \tau, \sigma)$
60 $If(S, P_1, P_2) \Rightarrow (\sigma_1 + \sigma_2), If^+(S, P_1^+, \sigma_1, P_2^+, \sigma_2)$
61 **where** $\sigma_1, P_1^+ = analyze(P_1, \tau, \sigma)$
62 **and** $\sigma_2, P_2^+ = analyze(P_2, \tau, \sigma)$
70 $Loop(P) \Rightarrow \sigma_2, Loop^+(P_2^+, (\sigma_2 - \sigma_1))$
71 **where** $\sigma_2, P_2^+ = analyze(P, \tau, \sigma_1)$
72 **where** $\sigma_1, P_1^+ = analyze(P, \tau, \emptyset)$
80 $Trap(T, P)\sigma \Rightarrow \sigma_1, Trap^+(T, P^+, \sigma)$
81 **where** $\sigma_1, P^+ = analyze(P, \tau[T \to \sigma], \sigma)$
90 $Exit(T) \Rightarrow \tau(T), Exit(T)$
100 $Signal(S, P) \Rightarrow \sigma_1, Signal^+(S, P^+, \sigma_1(S))$
100 **where** $\sigma_1, P^+ = analyze(P, \tau, \sigma[S \to 0])$

Figure 2: The prediction function

The main property of these predictions is that, even over-estimated, if a prediction assigns 0 to a signal, this signal is assured to be absent for the analysed instant.

The analysis decorates an Abstract Syntax Tree (AST). Figure 1 gives the correspondence between the syntax of core Esterel, the corresponding AST constructors and the decorated constructors. Note that we annotate only statements that need specific prediction adjustments at runtime.

The relation between the AST nodes and the decorated ones is defined in Figure 2. The function $analyze(P, \tau, \sigma)$ takes as arguments a node P, an environment τ mapping trap names to predictions, and the prediction σ for the continuation of P. The continuation of P is "*what remains to compute after the evaluation of P*". For a main program M, the analysis will be launched with $analyze(M, \bot_\tau, \emptyset)$,

where $\emptyset$ is the function that maps every signal name to zero, and $\perp_\tau$, is the empty environment. The function *analyze* computes both the prediction before P and P^+ which is the annotated version of P.

The prediction for Nop is trivial. The prediction for $Pause$ restarts the analysis, with an empty prediction $\emptyset$, for the instant before the $Pause$. The prediction of the continuation of the $Pause$ is saved in order to reset the prediction at the next instant. The rule for $Emit(S)$ increments by one the count for S. The rule for $Seq(P_1, P_2)$ is evaluated from right to left, and the result of P_2 in re-injected for P_1. The rules for $Par(P_1, P_2)$ and $If(S, P_1, P_2)$ are similar since they both sum the predictions of P_1 and P_2; the difference is that If saves these predictions in order to subtract one of them at runtime, while Par saves the prediction of its continuation; indeed, since at most one of P_1 or P_2 will execute this continuation, this continuation's prediction must be decremented at least once, twice when P_1 or P_2 raises an exception. The rule for $Loop(P)$ is more complex. Note that the prediction σ of the loop's continuation is useless since this continuation will never be activated. First, we have to analyse P with an empty prediction $\emptyset$; the result σ_1 is what we wait for the prediction of the continuation of P, thus we restart the analysis with this σ_1 prediction. If all paths traversing P cross a $Pause$[3], the final result σ_2 is equal to σ_1 and everything is said : no readjustment has to be done before reentering at the begin of the loop. Otherwise, if it exists at least one path traversing P without reaching a syntactic $Pause$, then σ_2 may differ from σ_1. Let's consider the simple example :

```
loop
      pause
   || emit S
end
```

The branch executing the **emit** S doesn't contain a **pause** but has to wait the other branch of the parallel; therefore, there is an implicit **pause** here. If we apply the rule of *Loop* given in Figure 2, the value of σ_1 for S is 1, and is 2 for σ_2. If we explicitly add a pause at the end of the second process, then both σ_1 and σ_2 have 1 for S. In all cases, a second tree traversal is needed to save correct predictions in the statements.

In case of instantaneous path (not crossing a pause) in a loop, σ_1 may be different to σ_2, but $\sigma_2 \geq \sigma_1$, thus $(\sigma_2 - \sigma_1)$ is a valid positive prediction to be added dynamically, at runtime, at the end of the loop body. This prediction readjustment is particularly non intuitive, hence we have decided to make a formal proof of the correctness of the analysis (see Section 5)

For the couple `trap`/`exit`, the prediction of the continuation of the trap is pushed on the environment and will be used by inner exits. This prediction is also saved in the AST node in order to make runtime adjustment as in the example presented in Section 2.

For local signals, we assume that there is no conflict between signal names, therefore we have only to save in the AST the prediction of the signal in order to set dynamically the counter before executing the body P. Due to the double evaluation of loops, the signal has to be reset before analysing the body. If it were not done, the following example, would compute a wrong value of 2 for the local signal S :

```
loop
   signal S in
         pause
      || emit S
end end
```

This is known as the "*reincarnation*" or the "*schizophrenia*" problem ([2] chap 12), it is significant only under the assumption that signals are mapped on wires of digital circuit and where the reactive machine cannot reset a signal during an instant. A cure of this problem can be found in [7], but, as we will see in the next section, an interpreter can easily reset the status of a signal and no duplication of code is needed.

[3]This is generally the case and we can start the analysis from this *barrier* of pauses to compute σ_1. In most cases, loop's bodies have to be analysed only once.

4 Evaluation

We have three main implementation decisions to take : first, we have to decide the style of evaluation, either by term rewriting (operational semantics) or by functional evaluation (denotational semantics). We choose the second style by describing the evaluation with a (may be recursive) function with explicit continuations (Continuation Passing Style). Continuations are adequate structures to denote processes and can have fast implementations. Next, we have to decide how to synchronise processes, i.e. how two processes launched in parallel, cooperate together to activate their common continuation. Finally, we have to decide how the set of processes is implemented, and more precisely how are managed the processes waiting for a signal status (present or absent).

4.1 CPS evaluation

All the running processes share a common information containing the state of signals. These shared values are stored in a structured data representing a *configuration*; we will take the variable μ for denoting these configurations. From a configuration μ and a signal S we can check that the signal is already emitted with the call $getEmit(\mu, S)$ and get the current count of emit prediction with $getCount(\mu, S)$. We can change these two values with a call to $ChangeSignal(\mu, S, emitted, n)$, where *emitted* is a boolean stating if the signal is currently emitted and n, its prediction counter.

The evaluation function given in Figure 3 accepts four parameters : P is the statement to be evaluated, γ is the stack of **parallel** and **trap** statements surrounding P, this stack is needed for the synchronisation and to store the continuation of a **trap**, μ is the global configuration containing the processes and all concerning signal states, and finally κ is the current continuation which is a function receiving a configuration as argument. For a main program M, we first run the static analysis with $analyze(M, \bot_\tau, \emptyset)$ and, as a result, we get an annotated statement M' and a prediction σ for the first instant. We create a first configuration μ with this prediction and we launch the evaluation with $eval(M', nil, \mu, \lambda\mu.\mu)$.

$eval(P, \gamma, \mu, \kappa) \stackrel{\triangle}{=}$ **match** P **with**
10 $Nop \Rightarrow \kappa(\mu)$
20 $Pause^+(\sigma) \Rightarrow doPause(\mu, \gamma, \sigma, \{\kappa\})$
$Emit(S) \Rightarrow$
30 $\kappa(changeSignal(\mu, S, true, getCount(\mu, S) - 1))$
40 $Seq(P_1, P_2) \Rightarrow eval(P_1, \gamma, \mu, \lambda\mu.eval(P_2, \gamma, \mu, \kappa))$
$Par^+(P_1, P_2, \sigma) \Rightarrow$
50 **let** $\gamma = Pslot(\sigma, \kappa) :: \gamma$ **in**
51 **let** $\kappa = \lambda\mu.doJoin(\mu, \gamma)$ **in**
52 **let** $\mu = addWork(\mu, \lambda\mu.eval(P_2, \gamma, \mu, \kappa))$ **in**
52 $eval(P_1, \gamma, \mu, \kappa)$
$If^+(S, P_1, \sigma_1, P_2, \sigma_2) \Rightarrow$
60 **match** $getEmit(\mu, S), getCount(\mu, S)$ **with**
61 **true**, $\_ \Rightarrow eval(P_1, \gamma, sigMinus(\mu, \sigma_2), \kappa)$
62 **false**, **0** $\Rightarrow eval(P_2, \gamma, sigMinus(\mu, \sigma_1), \kappa)$
63 $\_, \_ \Rightarrow stuck(S, \mu, \lambda\mu.eval(P, \gamma, \mu, \kappa))$
$Loop^+(P, \sigma) \Rightarrow$
70 **let** $t = getTime(\mu)$ **in**
71 **let** $\kappa = \lambda\mu.$**if** $c = getTime(\mu)$
72 **then** $Error$
73 **else** $eval(Loop^+(P, \sigma), \gamma, sigAdd(\mu, \sigma), \kappa)$
74 **in** $eval(P, \gamma, \mu, \kappa)$
80 $Trap^+(T, P, \sigma) \Rightarrow eval(P, Tslot(\sigma, \kappa, T) :: \gamma, \mu, \kappa)$
90 $Exit(T) \Rightarrow doExit(\mu, \gamma, T)$
$Signal^+(S, P, n) \Rightarrow$
100 $eval(P, \gamma, changeSignal(\mu, S, false, n), \kappa)$

Figure 3: The main evaluation function

The evaluation function makes a case analysis on the statement it has to evaluate. For the case of **nothing**, nothing have to be done, so the continuation is applied to the current configuration.

For a **pause** we have to do a synchronisation with the surrounding parallel statements. For this, we call a dedicated function *doPause* that will traverse the stack γ, the third argument σ is the prediction of emits for the current continuation. This prediction will be part of the global prediction when starting the next instant.

In case of an **emit** we change the state of the corresponding signal to be emitted. It is not necessary to decrement the global counter associated to S, but doing this decrement we can insure that when the counter reach a zero value, then no more emission of this signal can be done and therefore, in case of valued signals, all values are assigned to the signal.

Evaluating a sequence of statements in CPS is folklore; this is the main way of *pushing* a continuation. Evaluating the sequence of statements P_1 and P_2 with a continuation κ needs to evaluate P_1 with the continuation of evaluating P_2 with the original continuation κ.

The case of parallel statements is a more difficult. First, we push a new slot onto the stack saving the current continuation ant its prediction. These informations will be used when the parallel execution of P_1 and P_2 will try to synchronise. Then, we create a continuation which, when activated, will perform the synchronisation. After that, we create a new process whose behaviour is to evaluate P_2 in the new stack and with the join continuation. Finally, we continue by evaluating the left branch P_1. Note that a process is represented simply by a continuation and is added to configurations by the function $addWork$.

The test of the presence of a signal is the point where we use the global prediction. First, we test if the signal is already emitted; if so, we can directly evaluate the **then** branch with the same continuation. Since the other branch is not-taken we have to decrement the global prediction by the prediction for the not taken branch. $sigMinus(\mu, \sigma)$ creates a configuration similar to μ but with a global prediction decremented point-wise by σ. If the signal is not already emitted, we test its count. If this count has reached the value 0, it means that the signal cannot be emitted anymore for the current instant; thus, the signal is considered as absent and the **else** part is evaluated. When the signal is not emitted and its counter is strictly positive, then we cannot decide either if the signal will be emitted or, if it must be considered as absent. In this case, the process must wait until the status of the signal is defined. We will discuss several implementation of the *stuck* function in a next subsection.

For the **loop** statement, we have to evaluate the body in a dedicated continuation. First, this continuation has to check that the body is not instantaneously traversed in one single instant; for this we assume that the evaluation increments a logical clock every instant, the function $getTime$ returns the current clock value. Checking instantaneous loop is done by capturing the current clock value at the beginning of the loop and checking that the clock value is different at the end of the body. Once this check is done, we have to restart the entire loop while increasing the prediction by the amount computed by the static analysis.

For a **trap** statement, the evaluator has simply to evaluate the body by pushing a new slot onto the stack, saving the current continuation and its prediction. These saved values will be used by the synchronisation explained in a next subsection.

As for **pause**, an **exit** statement calls a dedicated function $doExit$ explained in a next subsection.

For a local signal declaration, we evaluate the body with a configuration where the signal is reset, i.e. the signal is declared as not emitted and its count is forced in the configuration to be the one computed by the static analysis; these two modifications are done by the function $changeSignal$. It is this ability to change the status of a signal, in the middle of an instant, which discards the *schizophrenia* problem. Moreover, the full renaming of local signals avoids to have to save and restore the counter of the signal at the entry and the exit of local definitions : the renaming insures that there is no other signal with the same name.

```
doExit(μ,γ,T) ≙ match γ with
10  Nil ⇒ EndOfProgram(μ)
20  Tslot(σ,κ,T1) :: γr ⇒ if T = T1
21                         then κ(μ)
22                         else doExit(μ,γr,T)
    Pslot(σ,κ) :: γr ⇒ match getState(μ,γ) with
      Started ⇒
30        schedule(changeState(μ,γ,Exited(T)))
40    Joined ⇒ doExit(μ,γr,T);
50    Paused(σ,l) ⇒ doExit(μ,γr,T);
      Exited(T1) ⇒
60        let T+,σ− = comp(γr,T,T1) in
61            doExit(sigMinus(μ,σ−),γr,T+);
```

Figure 4: Synchronization with an exit

```
doPause(μ,γ,σ,l) ≙ match γ with
10  Nil ⇒ EndOfInstant(μ,σ,l)
20  Tslot(σ,κ,T) :: γr ⇒ doPause(μ,γr,σ,l)
    Pslot(σ,κ) :: γr ⇒ match getState(μ,γ) with
      Started ⇒
30        schedule(changeState(μ,γ,Paused(σ,l)))
40    Joined ⇒ doPause(μ,γr,σ,l)
50    Paused(σ′,l′) ⇒ doPause(μ1,γr,σ+σ′,l • l′)
51        where μ1 = changeState(μ,γ,Started)
60    Exited(T) ⇒ doExit(μ,γr,T);
```

Figure 5: Synchronization with a pause

4.2 Synchronization

The three functions used in Figure 3, which do a process synchronisation, namely *doExit*, *doPause* and *doJoin*, receive as parameters a configuration and a stack. They perform a case analysis on the stack which can be empty or has on its top a slot left either by a **tag** or by a **parallel**. For a **parallel** slot, exactly two processes can make a synchronisation for this stack. The first incoming process will leave a trace on the stack, the last will make the final decision. For reading and writing this trace we use the couple of functions *getState* and *ChangeState*. The initial state of a parallel slot is assumed to be *Started* (implicit in the slot creation *Pslot*). Therefore, for these three functions, defined in figures 4, 5 and 6, finding a *Pslot* on top of the stack with a *Started* state (line 30 for all functions) will change this state to a dedicated one and stop the process : the function *schedule* takes a configuration as argument and restarts one suspended process.

As expressed in Figure 4, an *Exit* on an empty stack is dynamically checked even if it is refused at compile time. An $Exit(T)$ without an encompassing $Trap(T,\ldots)$ is considered as aborting the whole computation according to the proofs of the next section.

The objective of an *Exit* is to reach its corresponding *Trap* in the stack. There is one main case where unwinding the stack must be stopped : it occurs when a concurrent process wants also to exit (line 60 in fig 4), like in this pattern program :

trap T **in trap** T_1 **in** (**exit** T‖**exit** T_1) **end end**

The process executing "**exit** T" and the one executing "**exit** T_1" have the same stack, namely $Pslot(\sigma_1,\kappa_1) :: Tslot(\sigma_2,\kappa_2,T_1) :: Tslot(\sigma_3,\kappa_3,T) :: Nil$. The function *comp* returns the deepest tag name in the stack and the prediction associated to the lowest tag name. For the previous example, *comp* returns T and σ_2. Since the exit of the lowest tag (T_1) is not effective, we have to decrement its associated prediction (σ_2 or σ_- in Figure 4, line 61)

The *doPause* function (Figure 5) collects all the processes executing a **pause**. At a leaf of the tree of processes, when a process reaches a **pause** (fig 3, line 20), the function *doPause* is called with the continuation of the **pause** and the prediction for this process for the next instant. During the bottom-up traversal of the tree of processes (fig 5 line 50), the processes are appended in a list while the predictions are summed. One step of accumulation can be reflected by the transformation :

pause;P_1‖**pause**;P_2 ⇒ **pause**;(P_1‖P_2)

At line 50, the two sets of processes l and l' are still active (will continue to run in the next instant), therefore the *Pslot* on top of the stack is still alive and thus must be reset with a *Started* trace (line 51) for the next instant. When the collection of continuations ends at the root of the tree of processes (line 10), there is no more active process and we have reached the end of instant, i.e. all processes have

reached a **pause**. A new configuration can be made by incrementing the local time, setting the active processes with l and reseting the global prediction with σ. Note that when a process *restarts*, it retrieves *automatically* the stack it had when it was paused. This feature is due to the fact that stacks are closed under continuations. For example in fig 3 line 40, when evaluating the sequence $P_1;P_2$ in a stack γ, the evaluation of P_2 will be evaluated in the same stack.

The $Tslot$s on the stack are only there to be caught by a corresponding $Exit$; therefore both $doPause$ and $doJoin$ (fig 6) will pop it by doing a recursion on the rest of the stack (line 20 on fig 5 and 6). This can be understood by the transformations :

trap T **in pause**;P **end**
$\Rightarrow$ **pause**;**trap** T **in** P **end**

and

trap T **in** Nop **end** $\Rightarrow Nop$.

One interesting situation is when one branch of a **parallel** is paused and the other is finished, as in the pattern :

pause;$P \| $**nothing**

One can consider that the parallelism must vanish and the previous statement reduced to "**pause**;P". But the continuation of the paused process has closed the stack with the $Pslot$ on its top. It will be difficult and costly to remove this $Pslot$ in the interleaving of continuations. We will consider a lighter reduction of the previous statement with : "**pause**;($P \| $**nothing**)". The $Pslot$ will stay on the stack but its state will never be changed in further instants and will remain as $Joined$. Note that this $Joined$ state is forced when the paused process arrives first (fig 6 line 51) and otherwise is not changed (fig 5 line 40).

Recall (fig 2) that the prediction of the continuation of a parallel execution is *injected* in the two branches of the **parallel** statement. If these two branches terminate normally (i.e. doesn't issue an **exit**) the prediction of the continuation is counted twice. Therefore, the first branch that completes its execution has to decrease the global prediction. This is done in the $doJoin$ function when a finished process arrives first at a Pslot (fig 6 line 30)

Finally, the ultimate goal of the synchronisation is reached in fig 6 line 40 when two processes have completed the two branches of a **parallel** and when the continuation saved in the slot can be activated.

```
doJoin(μ,γ) ≜ match γ with
10  Nil ⇒ EndOfProgram(μ)
20  Tslot(σ,κ,T) :: γr ⇒ doJoin(μ,γr)
    Pslot(σ,κ) :: γr ⇒ match getState(μ,γ) with
30    Started ⇒ schedule(sigMinus(μ1,σ))
31      where μ1 = changeState(μ,γ,Joined)
40    Joined ⇒ κ(μ)
50    Paused(σ,l) ⇒ doPause(σ1,γr,σ,l)
51        where σ1 = sigMinus(μ1,σ)
51        where μ1 = changeState(μ,γ,Joined)
60    Exited(T) ⇒ doExit(sigMinus(μ,σ),γr,T);
```

Figure 6: Synchronization with a join

4.3 Waiting strategies

When a process reaches a "**present** S" statement where the signal is not emitted and doesn't have a count to 0, it has to wait (fig 3, line 63) via the *stuck* function. A first definition of this function could be simply :

$stuck(S,\mu,\kappa) \triangleq schedule(addWork(\mu,\kappa))$

It looks like a **busy-waiting** implementation. It seems better, and it is our current implementation, to have a specific set of waiting processes attached to signals. The *stuck* function simply adds the continuation to this set and proceeds with an other process via the *schedule* function. These waiting processes will be awaken as soon as the signal will be emitted, or when the counter of the signal becomes 0. This adds a test each time we emit a signal (fig 3 line 30) and an other test each time we decrement the count of a signal (inside of the $sigMinus$ function). In return, however, the set of processes managed by the configuration, updated via the $addWork$ function, contains only *active* continuations. Therefore when the *schedule* function is called with a configuration containing an empty set of active process, we can infer a *cyclic dependency* between the waiting processes, and raise an error in this circumstance. For example, the following program, which only emit a signal in

the continuation of its presence test, will raise such an error :

present S **then emit** S **else emit** S **end**

Some semantics of Esterel accept this program by considering the signal *emitted a priori* and by proving that considering the signal absent brings a contradiction. For this kind of program we assume that the user will rewrite the program by pulling up the **emit** before the test.

Coming back to the first definition of the *stuck* function, even if the waiting process is inserted again in the pool of processes, it may/must be the responsibility of the scheduler (i.e. the *schedule* function) to not pick up a new process arbitrarily. For example we can have two versions of the function $addWork$. The first version, used at process creation (fig 3 line 52), will add the process on the *left* of the pool of processes, as if this pool had a **LIFO**-*stack* structure. The second version, used in the *stuck* function, will add the process in the *right* of the pool as it had a **FIFO**-*queue* structure. The scheduler will use the pool as a stack when choosing a new candidate. This strategy doesn't add extra tests in **emit** or in $sigMinus$, but a waiting process can still be asked to compute. Consider the following pattern of program :

present S_1 **then emit** OK **end**
$\ldots \|$**present** S_{i+1} **then emit** S_i **end**
$\ldots \|$**emit** S_n

of the form $P_1 \| \ldots \| P_{n+1}$. After the first round all the processes but P_{n+1} are waiting for presence/absence of a signal and the last process P_{n+1} has emitted the signal S_n. At this point, the running list is $P_1 \ldots P_n$ and only P_n can react to the emission of S_n. Therefore this program will run in time proportional of n^2 whereas using the commutativity of the parallelism, the program $P_{n+1} \| \ldots \| P_1$ will run in a linear time. Moreover, with this implementation of waiting processes, it is not straightforward to decide when a *cyclic dependency* occurs.

4.4 Trampolines

Once written in CPS style, all significant calls of the evaluator are done in tail position. We can check in the figures 3 to 6 that the calls to *eval*, $doExit$, $doPause$, $doJoin$, *schedule* and all applications of continuation are in tail positions. Even the function $EndOfInstant$ can be written with a tail call to *schedule*, with the appropriate new configuration as argument. Therefore, the execution of all programs can be done with a fixed bound of system stack.

It may appear that the implementation language, in which the evaluator is written, consume some stack space even for tail calls. In this case a trampoline technique must be used. It consists to return an object representing a function call instead of applying it. This object can be simply a **thunk**, a.k.a a function without argument. As an example, one can decide to *clean* the system stack each time a **nothing** is evaluated. Even if is not the better place to do it, we can change the line 10 in Figure 3 by $\lambda\_.\kappa(\mu)$, this *frozen* computation is directly returned to the first call of the evaluator. The main function calls the evaluator with the main statement and, while the answer is not the one given by $EndOfProgram$, the frozen computation is restarted.

One can argue that adding a trampoline and finding places where to return thunks are implementation decisions and do not change the observation we can have on the computation : AST nodes are evaluated in the same order. But this equivalence helps to put in evidence some connections between different kinds of semantics. Without trampoline, the *eval* function defines a big step semantics of Esterel. If we return to the trampoline only inside the $EndOfInstant$ function, the *eval* function does the computation for one instant and is ready to be linked to standard operational semantics. If we return to the trampoline every function call (except for functions like $sigMinus$, $changeState$,...) we have switched to a small step semantics and the *eval* function, which is no more recursive, can be seen as defining a deterministic relation between configurations.

Since we have also specified this model of Esterel's computation with the system Coq and knowing that it is not natural to define not well founded recursion, i.e. without a proof or termination, it was a chance to have this opportunity to switch to the *equivalent* small-step definition.

5 Coq specification

From the previous presentation, we have introduced several simplifications in order to prove the correctness of the specification in Coq. (1) we have augmented the semantics with some traces in order to express the correctness. (2) we have specialised the general type of continuations (function from configuration to answer) to a dedicated inductive type enumerating the different kinds of continuations used by the semantics. This transformation is known as **defunctionalization** and described in [6] (3) we have unified the two passes prediction/evaluation by a single one where the evaluation recomputes, when needed, the prediction : a global prediction is no more used. (4) we have switched to a small step semantics where a step of computation is done by a (not recursive) function and the synchronisation is achieve by an inductive definition corresponding to a relation between configurations[4]. Finally we have proved the correctness of the specification in this simplified definition.

5.1 Domains

The traces that executions may leave are of three kinds :

```
Inductive event : Set :=
 | Emitted (s:name)
 | Present (s:name)
 | Absent (s:name).
```

At the end of an instant, a signal may be `Emitted`, considered as `Present` or considered as `Absent`. The correctness proof will insure that these three predicates are consistent, which mean essentially, that a message considered absent, with a counter valued to zero, must not be emitted . The evaluator will be tuned to generate `event`s to a trace added to configurations.

A configuration is a tuple containing the number of instants currently performed (`getTime`), the multi-set of processes (`getWork`) where a process is simply a continuation (a multi-set is needed since two processes may have the same continuation as in $P\|P$), the function remembering the emitted signals for the current instant (`getEmit`), the current trace of events (`getTrace`). The last slot of configurations (`getMax`) is used to generate fresh variables.

```
Record config := Config {
 getTime : nat;
 getWork : (list cont);
 getEmit : (name -> bool);
 getTrace : (list event)
 getMax : name;
}.
```

Then, for each explicit continuation expressed in Figure 3 (lines 40, 51, 52 and 71), for each direct recursive call to the evaluator in the same figure (lines 40, 53, 61, 62, 74, 80 and 100) and for each call to a synchronisation function (lines 20, 51 and 90) we create a dedicated inductive data structure.

```
Inductive cont : Set :=
 | KStep  (s:ast) (stk:(list slot)) (k:cont)
 | KLoop  (s:ast) (stk:(list slot)) (t:nat)
 | KExit  (stk:(list slot)) (t:name)
 | KPause (stk:(list slot)) (l:(list cont))
 | KJoin  (stk:(list slot)).
```

The domain of slots is also an inductive type reflecting the $Pslot$ and $Tslot$ expressed in the previous sections.

5.2 Specification

The main function of the specification is described in Figure 7 and is strongly related to the *eval* function given in Figure 3 receiving the same kind of parameters. Let's consider a configuration μ^+ whose slot $getWork$ is a list containing a continuation κ^+ of the form $KStep(P, \gamma, \kappa)$. Let's consider the configuration μ identical to μ^+ except for the slot $getWork$ where the continuation κ^+ is removed. In this state, the function *step* given in Figure 7 will be called with the arguments $step(P, \gamma, \mu, \kappa)$. Note that the configuration μ^+ is rebuild line 61. The behaviour of $step(P, \gamma, \mu, \kappa)$ is to compute a new continuation κ_1

[4]Since this relation doesn't introduce difficulties in the proofs, we will not describe it here.

$step(P,\gamma,\mu,\kappa) \stackrel{\triangle}{=}$ **match** P **with**

10 $Nop \Rightarrow \downarrow_{\mu}^{\kappa}$

20 $Pause \Rightarrow \downarrow_{\mu}^{KPause(\gamma,(\kappa::nil))}$

30 $Emit(S) \Rightarrow \downarrow_{addEvent(Emitted(S),changeSignal(\mu,S,true))}^{\kappa}$

40 $Seq(P_1,P_2) \Rightarrow \downarrow_{\mu}^{KStep(P_1,\gamma,KStep(P_2,\gamma,\kappa))}$

50 $Par(P_1,P_2) \Rightarrow \downarrow_{addWork(KStep(P_2,\gamma_1,KJoin(\gamma_1)),\mu)}^{KStep(P_1,\gamma_1,KJoin(\gamma_1))}$

51 **where** $\gamma_1 = Pslot(\kappa) :: \gamma$

60 $If(S,P_1,P_2) \Rightarrow$

61 **let** $\mu^+ = addWork(KStep(P,\gamma,\kappa),\mu)$ **in**

62 **match** $getEmit(\mu,S), getKCount(\mu^+,S)$ **with**

63 **true**, _ $\Rightarrow \downarrow_{addEvent(Present(S),\mu)}^{KStep(P_1,\gamma,\kappa)}$

64 **false**, **0** $\Rightarrow \downarrow_{addEvent(Absent(S),\mu)}^{KStep(P_2,\gamma,\kappa)}$

65 _, _ $\Rightarrow AStuck(\mu)$

70 $Loop(P) \Rightarrow \downarrow_{\mu}^{KStep(P,\gamma,KLoop(P,\gamma,getTime(\mu)))}$

80 $Trap(T,P) \Rightarrow \downarrow_{\mu}^{KStep(P,(Tslot(\kappa,T)::\gamma),\kappa)}$

90 $Exit(T) \Rightarrow \downarrow_{\mu}^{KExit(\gamma,T)}$

100 $Signal(S,P) \Rightarrow \downarrow_{\mu_1}^{KStep(rename(P,S,S_1),\gamma,\kappa)}$

101 **where** $S_1,\mu_1 = newSignal(\mu)$

Figure 7: simple evaluation

$\searrow_S (\kappa) \stackrel{\triangle}{=}$ **match** κ **with**

10 $KStep(P,\gamma,\kappa) \Rightarrow \stackrel{S}{\Pi}_{\Downarrow_S(\gamma)}^{\searrow_S(\kappa)} P$

20 $KLoop(P,\gamma,time) \Rightarrow \stackrel{S}{\Pi}_{\Downarrow_S(\gamma)}^{0} Loop(P)$

30 $KExit(\gamma,T) \Rightarrow cassq(T, \Downarrow_S (\gamma))$

40 $KPause(\gamma,l) \Rightarrow 0$

50 $KJoin(\gamma) \Rightarrow getJoinCounter(\gamma,S)$

Figure 8: continuations to counters

corresponding to the rest of the computation after one step of evaluation of P in a stack γ and a continuation κ; *step* also generates a new configuration μ_1 based on μ. Once κ_1 and μ_1 are build, the continuation is added in the working list of μ_1 and this final configuration is returned; this is depicted by $\downarrow_{\mu_1}^{\kappa_1}$.

We have added insertion of events (lines 30, 63 and 64). Note lines 100 and 101, that we do not consider a complete renaming of signals. Thus, local signals are managed with fresh variables dynamically generated.

The major difference, between the interpreter depicted in fig 3 and the one implemented in Coq, is that a global prediction is no more considered and thus the functions *sigMinus* and *sigAdd* (lines 61, 62 and 73 of Figure 3) are no more used. Each process, and each continuation, is individually responsible of its prediction. If, for a given signal S, we consider a function $\searrow_S (\kappa)$ from continuation to integer, computing the counter for S for this continuation, then for a configuration the global counter is the sum of the continuation based counters for the continuations found in the working list :

$$getKCount(\mu,S) \stackrel{\triangle}{=} \sum_{\kappa \in getWork(\mu)} \searrow_S (\kappa)$$

The correspondence between continuations and predictions is given in Figure 8. The crucial case if for a $KStep(P,\gamma,\kappa)$ continuation, which need to be related to $analyze(P,\tau,\sigma)$ of Figure 2. Here γ is a list of slots ($Pslot$ or $Tslot$) while τ is a mapping from tag names to predictions. It is easy to extract only $Tslot$ out of γ and, for each of them, to extract a counter via the continuation (using $\searrow$ ()) found in the $Tslot$; this stack projection is done with the

function $\Downarrow_S(\gamma)$. The counter for a $KJoin(\gamma)$ continuation is computed with the continuation found in the first $Pslot$ in γ, or returns 0 if no $Pslot$ is found in γ.

Since for a continuation κ we need the counter for a specific signal S ($\searrow_S(\kappa)$), we have to specialise the function $analyze(P,\tau,\sigma)$ for this signal. $\overset{S\ a}{\Pi}_\rho\ P$ computes the counter for S for the evaluation of P knowing that a is its counter for the continuation of P, ρ is a mapping from tag names to counters (implemented as list of associations for simplicity). Therefore, following Figure 2, we define $\overset{S\ a}{\Pi}_\rho\ P$ in Figure 9.

In this specification, we do not consider a complete renaming of the signals. It was simpler for the proofs not to characterise renamed programs. To avoid name conflicts, in Figure 7 line 100, each time a local signal definition is reached, a fresh signal name is generated. For the prediction (Figure 9 line 100), it is not trivial to going throw a local signal definition with the same name. In this case, the count of the continuation (i.e. a) seems correct but it would not consider the exceptions done by the body of the local definition. To be correct, this body must be analysed and the effect done on the local signal (i.e. $\overset{S\ 0}{\Pi}_{nil}\ P$) must be discarded from the result. The key point is the *nil* parameter which avoid to consider the access to S outside of P via an **exit**.

5.3 Correctness

We have to prove that the events added by the evaluator are coherent, that is, a signal considered present will be emitted in the same instant, and conversely a signal considered as absent will not be emitted in the same instant. A configuration is well formed ($WF(\mu)$) if it is reachable from an initial configuration build with a main statement.

Theorem 1 (correctness). *:* $\forall \mu, S,\ WF(\mu) \Rightarrow$
$[Present(S) \in getTrace(\mu)$
$\Rightarrow Emitted(S) \in getTrace(\mu)]$
$\wedge\ [Absent(S) \in getTrace(\mu)$
$\Rightarrow Emitted(S) \notin getTrace(\mu)]$

$\overset{S\ a}{\Pi}_\rho\ P \triangleq$ **match** P **with**
10 $Nop \Rightarrow a$
20 $Pause \Rightarrow 0$
30 $Emit(S_1) \Rightarrow$ **if** $S_1 = S$ **then** $a+1$ **else** a
40 $Seq(P_1,P_2) \Rightarrow \overset{s\ \overset{S\ a}{\Pi}_\rho P_2}{\Pi}_\rho\ P_1$
50 $Par(P_1,P_2) \Rightarrow \overset{s\ a}{\Pi}_\rho\ P_1 + \overset{S\ a}{\Pi}_\rho\ P_2$
60 $If(A,P_1,P_2) \Rightarrow \overset{s\ a}{\Pi}_\rho\ P_1 + \overset{S\ a}{\Pi}_\rho\ P_2$
70 $Loop(P) \Rightarrow \overset{S\ \overset{S\ 0}{\Pi}_\rho P}{\Pi}_\rho\ P$
80 $Trap(T,P) \Rightarrow \overset{S\ a}{\Pi}_{(T,a)::\rho}\ P$
90 $Exit(T) \Rightarrow cassq(T,\rho)$
100 $Signal(S_1,P) \Rightarrow \overset{S\ a}{\Pi}_\rho\ P - ((S_1 = S)\ ?\ \overset{S\ 0}{\Pi}_{nil}\ P : 0)$

Figure 9: simple prediction of a signal

Proof. The first part is proved with the lemma $getEmit(\mu,S) = true \Rightarrow Emitted(S) \in getTrace(\mu)$ which is not difficult since the *Emitted* event is added in the same step where the status of the signal is changed (line 20 of Figure 7). The second part uses a proof by contradiction using the following lemma. ∎

Lemma 1 (positive)**.** *A positive count is incompatible with an absent event:*
$\forall \mu, S,\ WF(\mu) \Rightarrow$
$getKCount(\mu,S) > 0 \Rightarrow Absent(S) \notin getTrace(\mu)$.

Proof. By induction on the step, induced by the WF definition, and by case analysis of the configuration for a computational step. All cases, except loop, are not difficult since *getKCount* is generally decreasing with time. The end of an instant is trivial as the traces are reset. The major difficulty is for a loop which needs to prove that $\overset{S\ 0}{\Pi}_\rho\ Loop(P) \geq \overset{S\ 0}{\Pi}_\rho\ Seq(P, Loop(P))$ which is done by contradiction using the following lemma. ∎

Lemma 2 (null count conservation by loops)**.** $\forall P, \rho, S,$

$$\overset{S\ 0}{\Pi_\rho}P \quad P = 0 \Rightarrow \overset{S\ \overset{S\ 0}{\Pi_\rho}P}{\Pi_\rho} \quad P = 0$$

Proof. This lemma cannot be proved directly by induction on P. We have to switch to a set of more general lemmas which have to be proved simultaneously and which describe the variation of $\overset{s\ a}{\Pi_\rho}\ P$ on ρ and a. Knowing that $\overset{s\ a}{\Pi_\rho}\ P = \overset{s\ 0}{\Pi_\rho}\ P + a * k$ for some k, we can apply this equality twice to prove our lemma. ❦

Lemma 3 (Derivative of prediction)**.**

$$\forall P, \rho, a, s,\ \overset{s\ a}{\Pi_\rho}\ P = \overset{s\ 0}{\Pi_\rho}\ P + a * \Delta_F(P)$$
$$\wedge\ \forall P, \rho, a, s,\ \overset{s\ a}{\Pi_{(T,c)::\rho}}\ P = \overset{s\ a}{\Pi_{del(T,\rho)}}\ P + c * \Delta_E(P, T)$$

Proof. By induction on P. The function *del* removes all instances of the tag T in the list ρ. Note that an **exit** T without a surrounding "**trap** T **in end**" is here considered as *jumping* to the end of the program. Even if this behaviour doesn't appear at runtime, for the proof we consider that, at line 90 of Figure 9, the default value of *cassq* is 0. The function Δ_F computes the number of control paths that go through its argument in the same instant (i.e. without crossing a pause), while Δ_E computes the number of times that the evaluation of its first argument may exit with the tag given as second argument. ❦

Here are the definitions of these two functions that are mutually recursive :

$\Delta_F(P) \stackrel{\triangle}{=}$ **match** P **with**
10 $Nop \Rightarrow 1$
20 $Pause \Rightarrow 0$
30 $Emit(sig) \Rightarrow 1$
40 $Seq(P_1, P_2) \Rightarrow \Delta_F(P_1) * \Delta_F(P_2)$
50 $Par(P_1, P_2) \Rightarrow \Delta_F(P_1) + \Delta_F(P_2)$
60 $If(A, P_1, P_2) \Rightarrow \Delta_F(P_1) + \Delta_F(P_2)$
70 $Loop(P) \Rightarrow 0$
80 $Trap(T, P) \Rightarrow \Delta_F(P) + \Delta_E(P, T)$
90 $Exit(T) \Rightarrow 0$
100 $Signal(sig, P) \Rightarrow \Delta_F(P)$

$\Delta_E(P, tag) \stackrel{\triangle}{=}$ **match** P **with**
10 $Nop \Rightarrow 0$
20 $Pause \Rightarrow 0$
30 $Emit(sig) \Rightarrow 0$
40 $Seq(P_1, P_2) \Rightarrow \Delta_E(P_1, T) + \Delta_F(P_1) * \Delta_E(P_2, T)$
50 $Par(P_1, P_2) \Rightarrow \Delta_E(P_1, tag) + \Delta_E(P_2, tag)$
60 $If(A, P_1, P_2) \Rightarrow \Delta_E(P_1, tag) + \Delta_E(P_2, tag)$
70 $Loop(P) \Rightarrow (\Delta_F(P) + 1) * \Delta_E(P, tag)$
80 $Trap(T, P) \Rightarrow (tag = T)?0 : \Delta_E(P, tag)$
90 $Exit(T) \Rightarrow (tag = T)?1 : 0$
100 $Signal(sig, P) \Rightarrow \Delta_E(P, tag)$

6 Related work and implementations

In this section we consider the prediction analysis and the evaluator traditionnaly used in the synchronous language approach (subsection 6.1) and then we will give some details of the implementation (subsection 6.2).

6.1 Position

Positioning Esterel in the synchronous model is largely discussed in Attar's thesis [1]. The closest area to Esterel is the reactive approach where the absence of signals is not predicted but simultaneously established when no process can make a step of calculus and the end of the instant is imposed. In this approach there is always a delay of one instant between the evaluation of a test and of its **else** branch (when taken). The implementation of this reactive approach is really simpler since predictions are not needed. If we consider an implementation where waiting processes are attached to signals (see the discussion in 4.3), the case where the scheduler is called with an empty set of active processes is no more a *cyclic dependency* but simply the end of the instant. In this case, all the waiting processes come back in an active state knowing that the specific signal **was** absent in the previous instant. This approach is very attractive, but there are still some programs that need to react without delay to absence. Think, for example, when one want to simulate a digital clocked circuit with each wire having its own signal. The presence of the signal acts as the hight level value in the wire,

the absence for the low level value. For example a negation for a wire *in* to a wire *out* is simply :

present *in* **then nop else emit** *out* **end**

Reacting instantaneously to the absence allows to maintain a direct relation between the clock of the circuit and the instants.

Using continuations for implementing/specifying synchronous languages is not widely adopted. The semantics is generally described by a structured operational semantics with term rewriting. The first use of continuations and of big step description seems to be attributed to L. Mandel [5] for the ReactiveML language. O. Tardieu and L. Mandel have also a, not (yet) published, interpreter, based on continuations, for a core of Esterel where the main ideas of our paper can be found. Instead of computing counters, list of lists of "accessible" emitters are extracted from the source. The contribution of our paper is to do this extraction at *compile* time.

In [4], F. Boussinot extracts, through potential functions definitions, a hierarchy of semantics with refinement of absence detection. It seems that our interpreter falls in the "v3" category.

6.2 Implementation

All the sources can be downloaded from ftp:ftp-sop.inria.fr/indes/rp/EsterelCounter.tar. The source code of the Coq specification is the file `correct.v` file. This file contains about 1500 lines with around 25% of definitions and 75% of proofs. The Scheme source code corresponding to the definitions found in Section 3 and Section 4 can be found in the file `esterelSpec.scm` and contains about 270 lines of code.

One remarkable characteristic of Esterel is the ability to compile a program into a digital circuit. This implies that all objects can be allocated statically. This is the reason why predictions can be computed statically, but in fact everything can be preallocated before running the evaluation : the stack, the continuations, the signals, the waiting list associated to signals, the running set of processes can actually be allocated statically. So we have implemented an interpreter able to run using a fixed size of memory. The Scheme code of this implementation can be found in the file `esterelStatic.scm` and contains about 830 line of code. Even if it is faster, this code is less readable than the previous version.

As stated in the beginning of the paper, Esterel is now candidate as an orchestration language with the implementation of HipHop [3] which is a layer of the **Hop** multi-tier language. We have also implemented a core of HipHop. The source code can be found in the file `hiphop.scm` which contains about 1600 lines of Scheme code. This code demonstrates that the extensions of HipHop can be handled with an implementation with counters, even with dynamic processes creation. The extensions proposed are :

1. **Values**. We have switched from statements to expressions, an expression returns a value. Any Scheme value can be introduced, as a constant, with the **quote** special form.

2. **Valued signals**. The **emit** special form accept an extra argument which is a value associated to the signal. The special form **present** is now the general Scheme test **if**. The presence/absence of a signal is tested with the special form **now** which return the true boolean (`#t`) when the signal is present, false otherwise. All the values emitted by a signal can be founded with the special form **val**. The counters associated to signals are used to insure that all the values of a signal are already emitted. When a signal is not emitted and have a counter to 0, then it is known to be absent, but also when a signal is emitted and have a counter to 0 then we known that no more value can be emitted.

3. **local variables**. We have introduced local variables declared with the **let** special form. The space need for these variables are preallocated and stored in the global memory.

4. **Scheme access**. All variables that cannot be resolved as a local one are considered as a Scheme global variable. All Scheme value accessible through a global variable can be accessed inside the reactive machine. This remains true for all global Scheme functions (+, string-ref, map...). We have also introduced a node dedicated to call a Scheme function. An expression

like `(+ 1 2)` is valid in the reactive machine, also the more interesting expression `(apply + (val sig))` which return the sum of all values emitted for the signal `sig`. The reactive machine is not accessible to Scheme, thus the evaluation of a Scheme code is instantaneous and cannot change the status of a signal. Evaluating the arguments of a function can take several instants, for example, knowing that the **pause** special form return the logical time of the reactive machine, the expression `(list (pause) 'quick (pause))` takes 3 instants and returns `(1 quick 2)`

5. **Dynamic process creation**. We have implemented the HipHop's special form `(mappar (lambda (x) P[x]) Vals)` where, for each value `x` of the list computed by `Vals`, a process `P`, dependant of `x`, is created. All these processes `P[x]` run in parallel, may be within different instants, and join all together before the `mappar` returns all the computed values. The difficulty is to compute the counters for such expression. Consider for example the expression `(par (val s) (mappar (lambda (x) (emit s x)) (iota 5)))`. How the first process computing `(val s)` will know that *all* values are emitted, as it depends on the length of the list computed by `(iota 5)`? Nevertheless, we can compute a common prediction σ for each process. In order to insure that at least one process is taken in account, the prediction σ is taken for the whole *mappar* expression, as if the computed list has a length of one. Dynamically, when the length n of the list is known, we readjust the global prediction with $(n-1) * \sigma$.

6. **Dynamic process creation** (cont). In the previous *mappar* special form, all the processes start in the same instant. We have implemented a more general special form `(control Binds Body)` where `Binds` is a list of bindings of the form `(name (lambda (x) P[x]))` where an expression dependant of a variable `x` is associated to a name. `Body` may contains an expression of the form `(detach (name Value))`. Informally, `control` acts as a pool of processes, initially containing only one process computing `Body`. Each time an expression `(detach (name E))` is executed, a new process evaluating `P[E]` is added to the pool of processes. All the processes do a global synchronisation and all the computed values are returned by the `control` special form. A prediction can be precomputed for each expression in the `Binds` of the `control`. Then each time a `detach` is reached by the static analysis, the precomputed prediction is added.

7 Conclusion

We have presented a static analysis to let helps the interpreter of Esterel programs to decide the instantaneous absence of a signal . The correctness of the analysis was formally proved. The analysis has been easily extended for a more functional language including dynamic process creation.

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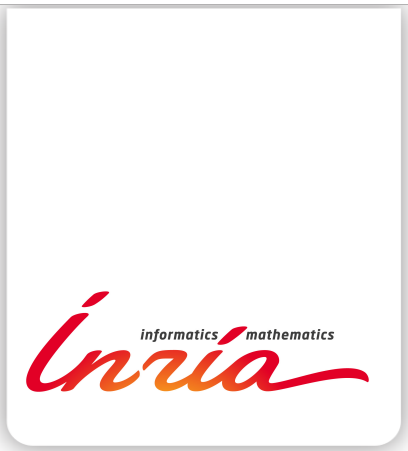

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