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A Hybrid Synchronous Language with Hierarchical Automata

Albert Benveniste
INRIA Rennes
albert.benveniste@inria.fr

Timothy Bourke
INRIA Paris-Rocquencourt
timothy.bourke@inria.fr

Benoît Caillaud
INRIA Rennes
benoit.caillaud@inria.fr

Marc Pouzet
DI, École normale supérieure
marc.pouzet@ens.fr

ABSTRACT

Hybrid modeling tools like Simulink have evolved from simulation platforms into development platforms on which testing, verification and code generation are also performed. It is critical to ensure that the results of simulation, compilation and verification are consistent. Synchronous languages have addressed these issues but only for discrete systems.

Reprising earlier work, we present a hybrid modeler built from a synchronous language and an off-the-shelf numerical solver. The main novelty is a language with hierarchical automata that can be arbitrarily mixed with data-flow and ordinary differential equations (ODEs). A type system statically ensures that discrete state changes are aligned with zero-crossing events and that the function passed to the numerical solver has no side-effects during integration. Well-typed programs are compiled by source-to-source translation into synchronous code which is then translated into sequential code using an existing synchronous language compiler.

Categories and Subject Descriptors

C.3 [Special-purpose and application-based systems]: Real-time and embedded systems; D.3.2 [Language classifications]: Data-flow languages

General Terms

Algorithms, Languages, Theory

Keywords

Real-time systems; Hybrid systems; Synchronous languages; Block diagrams; Compilation; Semantics; Type systems

1Located at DI, École normale supérieure.

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1http://www.mathworks.com/products/simulink

1. INTRODUCTION

Embedded systems are usually envisioned as discrete components that interact with a physical environment. This explains much of the success of hybrid modeling tools such as Simulink, which have progressively evolved from simulation platforms to development platforms where a single source code is used for simulation, testing, formal verification and the generation of embedded code. It is becoming critical to place such tools on a firm semantical basis to be able to prove that the results of simulation, compilation and verification are mutually consistent: simulation results must be producible; generated code must be faithful to the original model and efficient; verification results on the model should also be valid for generated code.

Synchronous languages [2] already addressed these issues for a class of discrete real-time systems for which time- and resource-bounded code can be generated. They abstract completely from the physical environment, treating it solely as a source of discrete sampled inputs and a recipient for discrete outputs. They concern themselves with the details of how discrete instants are produced but largely ignore the gaps in between. The physical environment, on the other hand, continues to evolve during such gaps. It is usually modeled with ordinary differential equations (ODEs) or with differential algebraic equations (DAEs) that interact closely with discrete controllers and may exhibit discontinuities at discrete instants. While a synchronous description of the whole is possible, replacing differential by difference equations, obtaining both efficiency and precision in simulations really necessitates the use of a numerical variable-step solver. This calls for a clear separation of continuous parts, which must be compiled specifically for activation by the solver, from discrete parts, which can be left unchanged.

Building on earlier work [1], we propose a language for hybrid systems by extending an existing synchronous language with ODEs and an off-the-shelf numerical solver (Sundials CVODE [8] is used in our prototype compiler). The synchronous language is used both for programming and as a target for code generation. The extension is conservative with respect to the synchronous subset: any synchronous function is compiled/optimized/executed exactly as if it were written in a regular synchronous language. There are two primary motivations for this approach. First,
synchronous languages are already a well-understood solution for modeling and compiling discrete components: their extension with hybrid features provides an integrated solution for modeling and simulating systems together with their environments. This is particularly relevant for systems that are mostly discrete, but that incorporate some physical elements. Second, when physical systems are simulated on a computer, it is not possible to observe continuous variables at all values of simulated time $t$. On the contrary, $t$ must be increased through a series of discrete increments. A synchronous language can thus act as a compilation target for hybrid models. Models are first translated into the synchronous subset of the language which is then compiled and linked with a numerical solver for the approximation of continuous trajectories. Note that the exact instants to sample are not necessarily known in advance nor periodic, but rather arise from the combined dynamics of continuous and discrete elements; that is, from the to and fro between a numerical solver and a discrete program. By translating the hybrid part into synchronous code and leaving the synchronous part unchanged, we can reuse most of the existing code generation and optimization techniques developed for synchronous languages [2] and also reduce the risk of differences between simulated and compiled code.

Contribution and organization of the paper.

The main novelty of this paper is a language which includes sophisticated control structures, namely hierarchical automata. These control structures can be composed in parallel or hierarchically, together with data-flow equations and ODEs with possible reset. Importantly, the control structures are valid in both discrete and continuous contexts. For example, the automata can be used to alternate between discrete control laws as in Mode Automata [12] or between sets of differential equations as in Hybrid Automata [7].

A type system is required to distinguish, at compile time, discrete computations from continuous ones. It ensures that discrete state changes are aligned with zero crossing events and determines which parts of a program are transformed for execution by a solver and which remain unchanged. Functions generated from well-typed programs are provably free of side effects when executed by a solver during the approximation of integral values. Maintaining this property is not easy when the bodies of control structures contain differential equations; there are several subtle typing issues. For example, transitions from one state to another in hybrid automata must only occur on zero-crossing events. One of the contributions of this paper is to clarify these issues.

We present a source-to-source transformation, for well-typed programs, that translates the full language into a synchronous subset amenable to code generation by a synchronous compiler. The ultimate goal is to translate programs written in a hybrid data-flow language with automata into imperative code, i.e., to go from the top-left of Figure 1 to the bottom-right. Given that synchronous programs with control structures (Data-flow+Auto) can already be translated to data-flow synchronous code (Data-flow) [12, 5], and that (Data-flow + ODE) can be translated to data-flow synchronous code (Data-flow) [1], there are two possible approaches. The first is to define a translation from (Data-flow+Auto+ODE) to (Data-flow) that removes control structures but not hybrid definitions. The second is to define a translation from (Data-flow+Auto+ODE) to (Data-flow+Auto) that removes hybrid constructs but not control structures. The latter translation, that effectively translates Hybrid Automata into Mode Automata, is considered in this paper. It is modular and maintains the structure of programming constructs.

This paper continues in Section 2 by recalling both the need for types when mixing discrete synchronous code with ODEs, and also the essentials of translating hybrid programs into synchronous code. It sketches the typing and compilation issues specific to control structures. In Section 3, the syntax of a basic synchronous language with control structures is presented, and the semantics of hierarchical automata are recalled. The extension of the language with ODEs is detailed in Section 4, and its type system is presented. The source-to-source translation is presented in Section 5. Related work is discussed in Section 6.

2. OVERVIEW

Following Lee and Zheng [10], we consider hybrid systems from a programming language perspective. We treat, in particular, the design, typing and compilation of a language which mixes data-flow equations and first-order ODEs. We do not consider verification.

Hybrid systems, by their very nature, involve the interwoven interaction of continuous dynamics and discrete events. Considered in isolation, a system’s continuous dynamics can be expressed as an initial value problem $x’ = f(t, x)$ init$x_0$, where $t \in \mathbb{R}_{\geq 0}$ represents physical time, $x$ is a vector of continuous state variables, $x’$ is a vector of instantaneous derivatives and $x_0$ is an initial state vector. Given $f$, a numerical solver (like SUNDIALS [8]) approximates the value of $x$ for increasing instants $t_1, t_2, \ldots$ of $t$. The gap between instants, the ‘step size’, may vary. And as, internally, a solver may jump back and forth when going from $t_i$ to $t_{i+1}$ to achieve acceptable accuracy, $f$ must be free of side effects, that is, combinational, at each (integration) step.

The first extension to this scheme to account for hybrid systems is to control the interruption of continuous trajectories. There are essentially two techniques: time horizons and vectors of zero-crossing expressions. Time horizons are useful for efficiently treating models that include periodically-executed discrete tasks. This is an important class of systems but not the only one: real applications also exhibit modes with discrete changes, in both the controller and the environment, that are not necessarily periodic or known in advance (e.g., an emergency braking system, or a clutch lock-up model, or a fuel control system)

2 See the SIMULINK/STATEFLOW automotive examples.
nism. Given a zero-crossing function \( z = g(t, x) \), a solver can check, during integration, for changes of sign in elements of \( z \) that may indicate discontinuities in continuous state variables, or other interesting events. Periodic timers can, in fact, be encoded as a particular form of zero-crossing. For example, this fragment defines a periodic signal:

\[
\text{der } p = 1.0 \text{ init } -2.0 \text{ reset } -2.0 \text{ every } z \text{ and } z = \uparrow(p)
\]

The value of \( p \) increases with slope 1 from an initial value of -2, and is reset to -2 every time \( p \) crosses zero.\(^3\) The variable \( z \) is true at the instants \((2n)_{n \in \mathbb{N}>0}\). Three types of action are possible in response to such discrete (zero-crossing) events: (1) reset elements of the continuous state vector \( x \) (as in the example above); (2) change the values of discrete state variables, which are implicit input parameters for \( f \) and \( g \); (3) trigger side-effects, e.g., update a display. There may be a sequence of such actions, a cascade, before the solver is invoked again to approximate continuous behaviors. Thus, the overall execution of a hybrid system alternates between continuous phases where integration is performed by a solver and discrete phases where side-effects may occur.

In \([1]\), an elementary, LUSTRE-like language extended with ODEs is introduced, along with a typing and compilation scheme for producing code linked with a black-box numerical solver. A type system is used to statically separate continuous and discrete execution phases where integration is performed by a solver and discrete execution phases where side-effects may occur. A signal is deemed discrete if it is specified on a discrete clock, defined \([1], [2]\):

\[
\text{A clock is termed discrete if it has been declared}
\]

so or if it is the result of a zero-crossing or a sub-sampling of a discrete clock. Otherwise, it is termed continuous.

A simple type system guarantees that when a program incorporates an ODE, all discrete changes, the aforementioned actions 1, 2 and 3, are aligned with zero-crossing events. The need for a type system can be understood from this example:

\[
\text{der time = 1.0 init 0.0 and } y = \text{sum}(time)
\]

In this example, an ODE defining \( \text{time}^4 \) occurs in parallel with a synchronous function \( \text{sum} \), whose definition is immaterial, but to be specific, could be assumed as:\(^5\)

\[
\text{let node sum}(x) = \text{cpt where}
\text{rec cpt = x -> (pre cpt +. x)}
\]

The example program is rejected because the computation of \( \text{time} \) is defined on a continuous clock while the computation of \( y \) is expected to run on a discrete clock.

Following \([1]\), we reject such programs statically. The intuition is to assign a kind \( k \in \{A, D, C\} \) to every expression and equation. An expression has kind \( D \) when it must be compiled with an internal discrete state and activated on a discrete clock, \( C \) when it must be compiled for activation by a solver on a continuous clock, and \( A \) when it is combinational and can be activated on any clock. In the example, \( \text{sum} \) is of kind \( D \), whereas \( \text{der } y = 1.0 \text{ init } 0.0 \) is of kind \( C \). The function \( \text{sum} \) cannot be activated directly on the stream \( \text{time} \) because the two have different kinds.

The original program should instead be written:

\[
\text{der time = 1.0 init 0.0 and}
\text{y = sum(time) every up(\text{ez}) init 0.0}
\]

where \( \text{ez} \) is an expression to be monitored for zero-crossing events. The semantics of \( y \) are now \( y(0) = 0 \), i.e., \( y \) is initialized to 0, and, if \( t_{a_1}, t_{a_2}, \ldots, t_{a_i}, \ldots \) are the instants of zero-crossing of \( \text{ez} \), \( y(t_{a_j}) = \sum_{i \leq j}(\text{time}(t_{a_i})) \), i.e., \( y \) is only computed at instants where \( \text{ez} \) crosses zero, and otherwise it keeps its previous value, i.e., \( y(t) = y(t_{a_j}) \) for \( t_{a_j} \leq t < t_{a_{j+1}} \).

Well-typed programs mixing ODEs and synchronous code activated on zero-crossings can be translated into purely synchronous code. Such a source-to-source translation ensures traceability between the source code and its operational form and allows the reuse of the existing optimization techniques in synchronous compilers. The principle is to translate a hybrid function activated on a continuous clock—one that contains ODEs and zero-crossings—into a regular synchronous function. The function is augmented with extra inputs for passing continuous state values and zero-crossing notifications, and also with extra outputs for returning instantaneous derivative values, updated continuous state values, and zero-crossing expression values. As an example, consider a bouncing ball with initial position \( y_0 \) and speed \( y'_0 \):

\[
\text{let hybrid bouncing}(y_0, y'_0) = y \text{ where}
\text{rec } dy = y' \text{ init } y_0
\text{and } dy' = -9.81 \text{ init } y'_0
\text{reset } -0.9 \cdot \text{last } y' \text{ every up(}-\text{.y)}
\]

Whenever the ball hits the ground (\( y \) changes from positive to negative), its speed is reset to \(-0.9 \cdot \text{last } y'\), where \( \text{last } y' \) is the left limit of signal \( y' \) when \( y' \) is left-continuous (when \( y' \) crosses zero), as is the case here, and the previous value of \( y' \) otherwise. The translation gives:

\[
\text{let node bouncing}((y_0, y'_0), ((\text{ly}, \text{ly'}), z)) =
(y, (\text{ly'}, ( \text{dy}, \text{dy'}), \text{upz})))
\text{where}
\text{rec } dy = y' \text{ and } y = y_0 -> \text{ly and upz = -.y}
\text{and } dy' = -9.81 \text{ and last_y'} = y'_0 -> \text{ly'}
\text{and } y' = \text{if x then -0.9 } \cdot \text{last_y'} \text{ else last_y'}
\]

The \( \text{der} \) definition of \( y \) is replaced with two new definitions: \( dy \) defines the instantaneous derivative of \( y \); and \( y \) is equal to \( y_0 \) at the very first instant, and thereafter to \( \text{ly} \), the value estimated by the solver.\(^6\) The \( \text{der} \) definition of \( y' \) is also replaced: initially \( \text{last_y'} \) is equal to \( y'_0 \), and thereafter it is equal to the input \( \text{ly'} \); \( y' \) takes the same values unless the zero-crossing input \( x \) is true, in which case it becomes \(-0.9 \cdot \text{last_y'}\). In general, an extra \( \text{last_y'} \) definition is needed to ensure correct initialization, in the presence of the \( \text{last} \) operator. The variable \( \text{upz} \) equals \(-.y\). In this system, \( \text{ly} \) and \( \text{ly'} \) are extra inputs to be computed by the solver whereas \( \text{dy} \) and \( \text{dy'} \) are extra outputs to be returned to the solver.\(^7\)

\(^3\)This can be written in Simulink as an integral with state port (here written \( st \)) and rising zero-crossing detection: \( p, st = \frac{1}{2}(-2.0, \uparrow(st), 1) \).
\(^4\)\( y \in \mathbb{R}^{\geq 0} \cdot \text{time}(t) = 0 + \int_0^t 1 \, dt = t \)
\(^5\)\( n \in \mathbb{N}, \text{sum}(x)(n) = \Sigma_{i=0}^n(x_i) \)
\(^6\)The LUSTRE initialization operator, \( \cdot -> \), has the semantics: \((x -> y)_n = x_0 \) and for all \( n \in \mathbb{N}^{>0} \), \((x -> y)_n = y_n \).
\(^7\)To avoid excessive copying, the extra inputs and outputs can instead be stored in arrays with in-place modification.
This translation is fine for a language without control structures. But control structures, such as hierarchical automata, are essential for designing complex systems with different operating modes. Such automata may appear at many levels of a system, in both discrete and continuous contexts. The problem, then, is not whether to extend the basic language, but rather how to do so.

We explain our approach with an example automaton:

```plaintext
let hybrid two(y0,y'0,z1,z2) = y where rec init y = y0 and automaton |
  Gravity -> local y' in do der y' = -9.81 init y'0 and der y = y' until z1 then Rising done |
  Rising -> do der y = 1.0 until z2 then Gravity done end
```

The initial value of \( y \) is given by the equation \( \text{init } y = y_0 \). The dynamics of \( y \) depend on the active mode of the automaton. In Gravity, the initial mode, \( y(t) = -9.81 \). In Rising, \( y(t) = 1 \). The notation until \( z_1 \) then Rising defines a weak preemption with reset: the body of state Gravity is executed before the condition \( z_1 \) is tested. If it is true, Rising becomes the active mode in the next reaction and the equations it contains will be reinitialized.

This program can be transformed into a discrete one.

```plaintext
let node two((y0,y'0,z1,z2), (ly,ly')) = (y, ((y,y'), (dy,dy'))) where rec y = y0 -> ly and automaton |
  Gravity -> do dy' = -9.81 and dy = y' and c = z until z1 then Rising done |
  Rising -> do dy = 1.0 and y' = ly' and dy' = 0.0 until z2 then Gravity done end
```

As before, two inputs, \( 1y \) and \( 1y' \), are added, as are two pairs of outputs: state variables \( y \) and \( y' \) and instantaneous derivatives \( dy \) and \( dy' \). While \( y' \) was initially declared local, it cannot be returned as an output unless its scope is extended, and it and its derivative must be given values in the other mode: we set them, respectively, to the estimate from the solver \( (1y') \) and to \( 0 \); inactive states are effectively suspended (the same convention as SIMULINK).

The result can be compiled by existing tools using standard techniques to transform automata into data-flow definitions [5].

In the previous example, the discrete events \( z_1 \) and \( z_2 \) are inputs. In general, discrete events may be computed locally and values may be communicated between states. Consider:

```plaintext
let hybrid bouncing(y0, y'0, start) = y where rec init y = y0 and automaton |
  Await -> do der y = 0.0 until start then Bounce(y'0) done |
  Bounce(y) -> local c in do dy' = -9.81 and dy = y' and c = z and upz = -. y and y' = v -> ly' until c & (y' < eps) then Await |
  | c then Bounce(-0.9 * . y') done |
```

The program has two modes. It waits in the initial one until the input start becomes true. It then enters the Bounce mode, which contains a variable \( c \) defined to be true when a zero-crossing occurs on the expression \(-. y\). When that event occurs, two transitions are possible. They are prioritized from top to bottom: if condition \( y' < \varepsilon \) is true (written \( c \text{ on } (y' < \varepsilon) \)), the automaton returns to the Await mode where the ball does not move; otherwise, Bounce is re-entered with a new value for the parameter \( v \), which is used in the concomitant reset of \( y' \) to model a bounce.

In the translation, the \( \text{der} \) definitions are treated as before. For the zero-crossing expression, a new input \( z \) replaces \( up(-. y) \), and a new output \( upz \) is defined to be \(-. y\).

```plaintext
let node bouncing((y0, y'0, start), ((ly, ly'), z)) = (y, ((y, y'), (dy, dy'), upz)) where rec y = y0 -> ly and automaton |
  Await -> do dy = 0.0 and dy' = 0.0 and upz = 0.0 and der y = ly' until start then Bounce(y'0) done |
  | Bounce(y) -> local c in do dy' = -9.81 and dy = y' and c = z and upz = -. y and y' = v -> ly' until c & (y' < eps) then Await |
  | c then Bounce(-0.9 * . y') done |
```

Transitions only occur on discrete events, viz. start and \( up(-. y) \); it is impossible for them to fire during integral approximation, as \( z \) and start will be false. The value of \( upz \) is only relevant if Bounce is active; it is set to \( 0 \) in Await.

Note that a discrete equation cannot be executed in a mode of a continuous automaton, even if, unlike the previous example with \( \text{sum} \), it does not interact with the other continuous variables. Consider, for example:

```plaintext
let hybrid wrong(z) = (y, c) where rec automaton |
  First -> do der y = 2.0 init 1.0 and c = 0 -> pre c + 1 until z then First done
```

In this automaton, an ODE \( y \) occurs in parallel with a discrete counter \( c \). If \([0,t_z] \subseteq \mathbb{R}^+ \) such that \( z(t_z) = \text{true} \) and for all \( t \in [0,t_z] \), \( z(t) = \text{false} \), initially, the semantics of \( y \) is:

\[
\forall t \in [0,t_z], y(t) = 1 + \int_0^t 2 \, dt = 1 + 2t
\]

The counter itself is defined on a discrete time basis. Then, what is the semantics of the parallel composition of the computation of \( y \) with that of \( c \)? Taking an initial part \([0..l] \subseteq \mathbb{N} \) for some \( l \in \mathbb{N} \), it could be tempting to state that:

\[
\forall n \in [1..l], c_n = c_{n-1} + 1 \text{ and } c_0 = 0
\]

But then, how should \([0,t_z] \) and \([0..l] \) be related and what is the value of \( l \)? The two equations do not share the same timescale (\( y \) runs on a continuous clock whereas \( c \) should be activated on a discrete one) and the meaning of their parallel composition is thus unclear. Moreover, once compiled, it would be necessary to continuously increment \( c \) during integration. This is an error which must be statically detected.\(^8\)

\(^8\)Rewritten in Stateflow, the compiler complains: In continuous time Stateflow charts, outputs cannot be read if they
Perhaps surprisingly, this new case is subsumed by the previously proposed system of kinds, which rejects \textbf{wrong}. The system of kinds clarifies some of the restrictions imposed by the MathWorks \textsc{Stateflow} compiler.

The source-to-source translation of automata is only valid if the generated transition and zero-crossing functions are side-effect free. This is guaranteed by extending the analysis that assigns kinds. Basically, a kind is assigned to a whole automaton, and the equations in the bodies of every state must have that kind. It is possible, though, in automata of kind \texttt{C}, to put discrete computations on transitions since they can only be activated on a discrete clock. This simple type discipline ensures that, once compiled, the transition function is free of side-effects during integration. In the next two sections, we formalize the static typing of programs with hierarchical automata and the source-to-source translations.

3. A SYNCHRONOUS LANGUAGE WITH AUTOMATA

3.1 Syntax

We first consider a synchronous language that mixes data-flow equations and hierarchical automata. Hybrid features are added in Section 4.

\begin{align*}
    d &::= \text{let } k \ f(p) = e \mid d \ d \\
    k &::= \text{D} \mid \text{A} \\
    e &::= x \mid v \mid \text{efby} e \mid \text{last}(x) \mid f(e) \mid (e, e) \mid \text{let E in } e \\
    p &::= (p, p) \mid x \\
    E &::= x = e \mid E \text{ and } E \\
    &\quad \mid \text{automaton} (S(p) \rightarrow u \text{ unless } s^*)^+ \\
    u &::= \text{local } x \text{ in } u \mid \text{do } E \text{ until } s^* \\
    s &::= e \text{ then } S(e) \mid e \text{ continue } S(e)
\end{align*}

A program is a sequence of global declarations (\(d\)) of functions over signals (\(\text{let } k \ f(p) = e\)), where \(k\) is the kind of the function: \(k = \text{A}\) declares that \(f\) is combinatorial; \(k = \text{D}\) declares that \(f\) may contain delays or automata.\footnote{The keywords \texttt{unless} and \texttt{until} are omitted in the concrete syntax for empty transition sequences.} Expressions are composed of variables (\(x\)), immediate values (\(v\)), initialized delays (\(\text{efby } e\)), last values of signals (\(\text{last}(x)\)), applications of functions to arguments (\(f(e)\)), pairs (\(e, e\)) and local declarations (\(\text{let E in } e\)), which ultimately take the value of an \(e\) that may contain variables defined in a set of equations \(E\). Patterns, \(p\), are variables (\(x\)) or pairs of patterns (\(p, p\)). Other tuples are shorthand for nested pairs.

Equations are denoted \(E\) and include variables \(x\) of value \(e\) (\(x = e\)), sets of equations in parallel (\(E \text{ and } E\)), local declarations of variables \(x\) in equations (\(\text{local } x \text{ in } E\)), and automata (\(\text{automaton} (S(p) \rightarrow u \text{ unless } s^*)^+\)).

An automaton is a list of handlers (\(S(p) \rightarrow u \text{ unless } s^*\)): \(S(p)\) is a parameterized state name, \(u\) is the state body and \(s^*\) is a sequence of strong transitions. The body \(u\) may contain variables local to the state (\(\text{local } x \text{ in } u\)) but it is ultimately a set of equations \(E\) paired with a sequence of weak transitions (\(\text{do } E \text{ until } s^*\)).\footnote{The keywords \texttt{unless} and \texttt{until} are omitted in the concrete syntax for empty transition sequences.} There are two types of transitions: either the target state is reset on entry (\(e \text{ then } S(e)\)) or it is resumed (\(e \text{ continue } S(e)\)), i.e. entered with history. \(S(e)\) gives the the target state name (\(S\)) and a value (\(e\)) to pattern match against that state’s parameters. The shorthand \(S\), i.e., a state with no parameters, stands for \(S(x)\) in equations, with \(x\) a fresh variable, and \(S(()\)) in transitions.

3.2 Examples and Intuitive Semantics

The distinction between combinatorial (kind \texttt{A}) and stateful functions (kind \texttt{D}) is necessary as they have different implementations. Combinatorial functions are typically imported from the host language (e.g., integer or boolean operators) and have no internal state. Such functions can be composed to form new combinatorial functions which are translated directly into the host language. Their type signature is of the form \(t_1 \overset{\text{A}}{\rightarrow} t_2\) with \(t_1\) and \(t_2\) as input and output types respectively. A stateful function, in comparison, may contain stateful operators (e.g., \texttt{fby, pre, \rightarrow}), automata or other stateful functions. Their implementation in the host language needs an extra argument for passing a mutable internal state record—all instances of such a function share the same compiled code but each is allocated a distinct static value. Their type signature is of the form \(t_1 \overset{\text{D}}{\rightarrow} t_2\). The kind of a function is validated by static typing.

The semantics of data-flow equations is the one of Lustre and not recalled here. Automata deserve more attention. We adopt the automata of Lucid Synchroné [14], whose detailed semantics is given in [4]. The two transition kinds (strong and weak) and the two ways of entering a target state (history and reset) give four possible types of transitions, which is typical in the majority of languages or tools for embedded system design that have hierarchical automata, like, for example, the hierarchical automata of SCADE 6, Ptolemy II [9], Statecharts [6], and Stateflow. An automaton’s behavior in a reaction depends on its immediate prior state \(S_{\text{prior}}\); initially the first in the sequence (which must take \(\emptyset\) as argument). It is computed in four steps:

1. For strong preemptions, we consider the \((s_i)_{i \in I}\) of the handler \(\text{unless}(s_i)_{i \in J} \overset{\text{D}}{\rightarrow} S_{\text{prior}}\). Each \(s_i\) has either the form \(e_i \text{ then } S_i(s_e_i)\), or the form \(e_i \text{ continue } S_i(s_e_i)\). All \((e_i)_{i \in I}\) are computed to give a list of booleans.

2. If none of the \((e_i)_{i \in I}\) are true, the current state remains active for the reaction: \(S_{\text{active}} = S_{\text{prior}}\). Otherwise, the branch corresponding to the earliest true expression \(e_k\) gives the active state: \(S_{\text{active}} = S_k(v)\), where \(v\) results from evaluating \(s_e_k\).

3. The body \(u\) of the active state \(S_{\text{active}}\) is reset if the last transition, be it strong or weak, had the form \(\cdot \text{ then } \cdot\). Execution passes through all \(u = \text{local } x \text{ in } u'\) into \(u'\), and when \(u = \text{do } E \text{ until } (s_j)_{i \in J}\).\footnote{In the concrete syntax, \texttt{A} is omitted and \texttt{D} is written \texttt{node}.} \(E\) is executed. Afterward, all weak transitions \((s_j)_{i \in J}\) are computed to find the state in which to begin the next reaction. Each \(s_j\) has one of the forms \(e_j \text{ then } s_e_j\), or \(e_j \text{ continue } s_e_j\). All \((e_j)_{i \in J}\) are computed.

4. If none of the \((e_j)_{i \in J}\) are true, the next state, i.e. \(S_{\text{prior}}\) at the next reaction, will be \(S_{\text{active}}\). Otherwise, the branch corresponding to the earliest true expression \(e_k\) gives the next state \(S_{\text{active}}(v)\), where \(v\) results from immediately evaluating \(s_e_k\).
4. HYBRID SYNCHRONOUS LANGUAGE

This basic language is extended with three constructs: ordinary differential equations, tests for defining zero-crossing events and conditionals on those events. We present a minimal extension which neglects ODE resets and synchronous code activations on zero-crossings. They are, however, considered elsewhere [1] and implemented in our prototype.

4.1 Extended Syntax of the Language

In this section, we consider an extended language:

\[
\begin{align*}
  k & ::= D \mid A \mid C \\
  e & ::= \ldots \mid up(e) \mid e \text{ on } e \\
  E & ::= \ldots \mid init x = e \mid \text{der } x = e \\
\end{align*}
\]

A new element \( C \) is added to the set of kinds \((k)\), for functions whose bodies must be activated on a continuous clock.

There are two new expression types. \( up(e) \) detects a rising zero-crossing on \( e \), i.e., the instant when \( e \) changes from negative to non-negative and \( e_1 \) on \( e_2 \) is present if \( e_1 \), a zero-crossing, is present and \( e_2 \), a boolean expression, is true.

The equations are extended with declarations of an initial value \( e_0 \) for \( x \) \((\text{init } x = e_0)\), and an instantaneous derivative \((\text{der } x = e)\). The two are used in combination to define an ODE \( x' = e \) with initial value \( e_0 \), for which we also allow the syntactic shortcut: \( \text{der } x = e \text{ init } e_0 \).

4.2 Static Typing

As discussed in Section 2, we must distinguish at compile time what is discrete from what is continuous. We adopt the convention that a signal is typed discrete if it is declared so or is activated on a zero-crossing event using the result of an \( up(\cdot) \) or \( \cdot \text{ on } \cdot \) construct. Otherwise, it is typed continuous.

Instead of associating this information with signals, we take a simpler approach and associate it with functions: a function \( f \) is given a type signature \( t_1 \rightarrow t_2 \) where \( t_1 \) is the type of its input, \( t_2 \) is the type of its output and \( k \) is a kind. If \( k = C \), \( f \) may only be used in a continuous context. If \( k = D \), \( f \) may only be used in a discrete context. If \( k = A \), then \( f \) is a combinatorial function that can be used in expressions of any kind. There is a subkind relation where, for all \( k \), \( k \subseteq k \) and \( A \subseteq k \).

The type language is:

\[
\begin{align*}
  \sigma & ::= \forall \beta_1, \ldots, \beta_n. t \rightarrow t \\
  t & ::= t \times t \mid \beta \mid bt \\
  k & ::= D \mid C \mid A \\
  bt & ::= \text{unit} \mid \text{float} \mid \text{int} \mid \text{bool} \mid \text{zero}
\end{align*}
\]

where \( \sigma \) defines types schemes and \( \beta_1, \ldots, \beta_n \) are type variables. A type \( t \) is either a pair \((t \times t)\), a type variable \( \beta \) or a base type \( bt \), and \( \text{zero} \) is the type of zero-crossing conditions, with \( up(\cdot) \) and \( \cdot \text{ on } \cdot \) as the only constructors.

As is typical, the typing rules are defined with respect to typing environments. We write \( G \) for the environment that accumulates over the sequence of global declarations in a program, mapping each to a type scheme \((\sigma)\). And we write \( H \) for a local environment mapping each variable \( x \) to its type \( t \) and status: \( \text{init}(x) : t \) if the initial value is defined but not the derivative, \( \text{der}(x) : t \) if the derivative is defined but not the initial value, \( \text{last}(x) : t \) if the last(\cdot) operator may be applied, and \( x : t \) otherwise.

\[
\begin{align*}
  G & ::= [f_1 : \sigma_1, \ldots, f_n : \sigma_n] \\
  H & ::= [] \mid H, x : t \mid H, \text{last}(x) : t \mid H, \text{init}(x) \mid H, \text{der}(x)
\end{align*}
\]

Operations on Environments.

If \( H_1 \) and \( H_2 \) are two environments, \( H_1 + H_2 \) is their union, provided their domains are disjoint; \( H_1, H_2 \) is their concatenation; \( H_1 + H_2 \) is a new environment such that

\[
\begin{align*}
  (H_1 + [x : t]) & \times (H_2 + [x : t]) = (H_1 + H_2) + [x : t] \\
  (H_1 + \text{last}(x) : t) & \times (H_2 + \text{last}(x) : t) = \\
  (H_1 + H_2) & + [\text{last}(x) : t] \\
  (H_1 + \text{init}(x) : t) & \times (H_2 + \text{der}(x) : t) = \\
  (H_1 + H_2) & + [\text{last}(x) : t]
\end{align*}
\]

(whence + and \( \times \) are associative and commutative). The initial environment \( G_0 \) gives the type signatures of imported operators (of kind \( A \)), synchronous primitives (of kind \( D \)), and zero-crossing constructors:

\[
\begin{align*}
  (+) & : \text{int} \times \text{int} \rightarrow \text{int} \\
  (=) & : \forall \beta, \beta \rightarrow \beta \\
  \text{pre}(\cdot) & : \forall \beta, \beta \rightarrow \beta \\
  \beta \rightarrow \beta & : \forall \beta, \beta \rightarrow \beta \\
  \text{up}(\cdot) & : \text{float} \rightarrow \text{zero} \\
  \cdot \text{ on } \cdot & : \text{zero} \rightarrow \text{bool} \rightarrow \text{zero}
\end{align*}
\]

Remark 1 \((up(\cdot))\) \(REAL\) or \(BOOLEAN\)?). The function \( up(x) \) takes an \( x \) of type \text{float}. Alternatively, \( x \) could be a \text{bool}; with the mapping \( \{t \rightarrow 1.0, f \rightarrow -1.0\} \) for the solver. The methods are equally expressive, but the modified second method in SUNDIALS converges faster in the first.

The type of transition guards.

Transition guards in discrete automata have type \text{bool}. They are evaluated when a reaction occurs and the source state of their transition is active. Guards in continuous automata are the same conceptually, but in practice discrete events are communicated as zero-crossings. Thus, we write \( k(\text{bool}) \) for the type of guards in context \( k \), which is defined \( C(\text{bool}) = \text{zero}, D(\text{bool}) = \text{bool} \) and \( A(\text{bool}) = \text{bool} \).

Generalization and Instantiation.

The types of program declarations are generalized to type schemes \((\sigma)\) by quantifying over all free variables, which is valid since they are defined globally.

\[
\begin{align*}
  \text{gen}(t_1 \rightarrow t_2) & = \forall \beta_1, \ldots, \beta_n. t_1 \rightarrow t_2 \\
  \text{if } \{\beta_1, \ldots, \beta_n\} & = \text{FV}(t_1 \rightarrow t_2)
\end{align*}
\]

The variables in a type scheme \( \sigma \) can be instantiated, and the kind \( k \) of a function can be replaced with any kind \( k' \) where \( k \subseteq k' \). Let \( \text{Inst}(\sigma) \) be the set of all such instantiations.

\[
k \subseteq k'
\]

\( (t \rightarrow t')[t_1/\beta_1, \ldots, t_n/\beta_n] \in \text{Inst}(\forall \beta_1, \ldots, \beta_n. t \rightarrow t') \)

Typing is defined by the predicates \((\text{typ-env})\) stating that the equation \( E \) produces the type environment \( H' \) and has kind \( k \), \((\text{typ-exp})\) stating that, in the global environment \( G \) and the local environment \( H \), an expression \( e \) has type \( t \) and kind \( k \), and \((\text{typ-pat})\) which defines the type and environment produced by a pattern \( p \), for which kinds are unnecessary.

\[
\begin{align*}
  (\text{typ-env}) & \quad (\text{typ-exp}) & \quad (\text{typ-pat}) \\
  G, H \vdash_k E : H' & \quad G, H \vdash_k e : t & \quad \vdash_k p : t, H
\end{align*}
\]
Three other predicates are used for typing automata: \( \text{(until)} \) defines the environment defined by the body and computes the list of target states together with the types of their parameters; \( \text{(then-continue)} \) defines the target state which can be reached by a handler \( e \) \( \text{then} \) \( se \) or \( e \) \( \text{continue} \) \( se \), \( k_1 \) is a restriction on the kind of the guard expression, and \( k_2 \) gives a restriction on its type; and \( \text{(s-param)} \) returns the type and environment associated with a state.

\[
\begin{align*}
\text{(INIT)} & \quad G, H \vdash e : t \\
\text{(DER)} & \quad G, H \vdash e : \text{float} \\
\text{(AND)} & \quad G, H \vdash E_1 : H_1 \quad G, H \vdash E_2 : H_2 \\
\text{(EQ)} & \quad G, H \vdash e : t \\
\text{(APP)} & \quad t \downarrow t' \in \text{Inst}(G(f)) \quad G, H \vdash e : t \\
\text{(LETIN)} & \quad G, H, H_0 \vdash k \ E : H_0 \quad G, H, H_0 \vdash e : t \\
\text{(const)} & \quad G, H \vdash k_4 : \text{int} \\
\text{(last)} & \quad G, H \vdash \text{last}(x) : t \\
\text{(VAR)} & \quad G, H \vdash [x : t] : t \\
\text{(VAR-LAST)} & \quad G, H \vdash [\text{last}(x) : t] : t \\
\text{(PAIR)} & \quad G, H \vdash k_1 : t_1 \quad H \vdash k_2 : t_2 \\
\text{(PAT-PAIR)} & \quad \vdash k_1 p_1 : t_1, H_1 \quad \vdash k_2 p_2 : t_2, H_2 \\
\text{(PAT-SEQ)} & \quad \vdash k_1 p_1 : t_1 \times t_2, H_1 + H_2 \\
\text{(DEF-SEQ)} & \quad \vdash \text{let } k \ f(p) = e : [f : \text{gen}_G(t_1 \rightarrow t_2)] \\
\text{(auto)} & \quad \forall i \in I, j \in J_i, \vdash \text{sp} : [S_i : t_i], H_i \quad G, H + H_1 \vdash u_i : H'_i, [S_i : t_i]_{i \in L_i} \quad G, H + H_1 \vdash s_j : [S_j : t_j] \\
\text{(local-in)} & \quad G, H \vdash k \ \text{local } x \ \text{in } u : H' \\
\text{(DO-UNTIL)} & \quad G, H \vdash k \ E : H' \quad \forall j \in J \quad G, H \vdash s_j : [S_j : t_j] \\
\text{(PARAM-STATE)} & \quad \vdash p : t_p, H_p \\
\text{(then)} & \quad G, H \vdash k_1 : e : k_2(\text{bool}) \quad G, H \vdash \text{do } E \ \text{until} (s_j)_{j \in J} : H', [S_j : t_j]_{j \in J} \\
\text{(continue)} & \quad G, H \vdash k_1 : e : k_2(\text{bool}) \quad G, H \vdash \text{continue } S(se) : [S : t] \\
\text{(s-param)} & \quad \vdash S(p) : [S : t_p], H_p \\
\end{align*}
\]

The typing rules are presented in Figures 2 and 3. Figure 2 is treats data-flow primitives.

Rule \( \text{(INIT)} \). An initialization of a variable \( x \) is well-typed if the defining expression \( e \) is well-typed. It takes the type \( t \) and kind \( k \) of \( e \). The status of \( x \) is marked \( \text{init}(x) \).

Rule \( \text{(DER)} \). An expression \( e \) defining the derivative of a variable \( x \) must be of type \( \text{float} \) and kind \( C \), and there must be an accompanying \( \text{init} \). The status of \( x \) is marked \( \text{der}(x) \).

Rule \( \text{(AND)} \). Parallel equations must be well-typed with the same kind.

Rule \( \text{(EQ)} \). The kind \( k \) and type \( t \) of a variable \( x \) defined by an equation are those of the expression \( e \), provided it is well-typed.

Rule \( \text{(APP)} \). A function application is well-typed if the kind \( k \) and argument type \( t \) of the function \( f \) can be instantiated to match those of a well-typed argument \( e \). The result has (instantiated) type \( t' \) and the same kind.

Rule \( \text{(LETIN)} \). A local definition is well-typed if the equations \( E \) and resulting expression \( e \) are well-typed in an extended environment. Both must have the same kind \( k \).

Rule \( \text{(const)} \). Set of rules illustrated by an integer constant.

Rules \( \text{(last)} \), \( \text{(VAR)} \) and \( \text{(VAR-LAST)} \). Occurrences of \( \text{last}() \) and variables must be respect the enclosing environment.

Rule \( \text{(PAIR)} \). The rule for pairing is similar to the one for the parallel composition of equations.

Rules \( \text{(PAT-PAIR)} \) and \( \text{(PAT-VAR)} \). These rules build an initial environment for patterns.
Rule (def-node). The type of a node is generalized from the types of the input pattern \( t_1 \) and defining expression \( t_2 \). The kind \( k \) of the expression \( e \) must match the declared kind.

Rule (def-seq). Function definitions are typed sequentially.

Figure 3 defines the typing of control structures:

Rule (auto). In the typing rule for automata, the kind \( k \in \{D,C\} \). The state pattern of each handler \( sp_i \) is typed, associating the state label \( S_i \) with the type of its parameter, and producing an environment \( H_i \), comprising the variables bound in the pattern and their types. The body and weak preemption conditions \( u_i \) are typed in a combination of the global environment \( H \) and that of the parameter variables \( H_i \) to yield an environment \( H_i' \), comprising the names defined by the body, and, for each condition, the destination state label \( S_j \) and the type of the associated expression \( t_j \). The strong preemption conditions \( s_j \) are typed similarly. The (auto) rule is conditional on additional constraints, stating that target states of strong and weak preemptions belong to the set of states defined by the automaton:

\[ [S_t : t_i]_{i \in I} \subseteq [S_t : t_i]_{i \in I} \quad [S_j : t_j]_{j \in I} \subseteq [S_t : t_i]_{i \in I} \]

Rule (local-in). An \( x \) defined in \( u \) can be made local to \( u \).

Rules (do-until) and (param-state). These rules apply respectively to state bodies \( E \) (and their weak preemption conditions \( s_j \)) and to state patterns.

Rules (then) and (continue). Guard expressions \( e \) must be of type \( k_2(\text{bool}) \), which is \( \text{bool} \) for discrete automata and \( \text{zero} \) for continuous ones. This constraint guarantees that transitions are only taken in response to discrete events. The context of the guard expression is given by \( k_1 \). For strong transitions, it will be \( A \) (combinatorial), and for weak transitions it will either be \( C \) or \( D \) depending on the overall kind of an automaton. The expression yielding the argument for the destination state \( se \) is always of kind \( D \), which allows synchronous functions to be executed in continuous automata when a transition fires.

5. TRANSLATION TO SUBSET

Well-typed programs can be translated to the synchronous subset of the language. The translation removes ODEs and zero-crossings and leaves discrete computations unchanged. It yields a sequential transition function that can be processed by a synchronous compiler.

The translation replaces zero-crossing operators and ODE definitions. Each zero-crossing operator \( up(e) \) is replaced by adding a new input \( z_i \) of type \( \text{bool} \), which replaces the operator occurrence, and a new output \( u_i \) of type \( \text{float} \times \text{bool} \) with a defining equation. For each ODE \( \text{der} x = e \) with initialization \( \text{init} x = e_0 \) we add a new input variable \( ly_i \), two new output variables \( x \) and \( dx \), and a local variable \( last_x \). Two new equations are added to define \( dx \) and \( last_x \), and the original equation defining \( x \) is modified.

Extra inputs and outputs are collected into vectors during the translation. For the inputs, we write \( v \) to denote a vector of zero-crossing variables \( [z_1, \ldots, z_n] \), and \( lxv \) for a vector of continuous state variables \( [x_1, \ldots, x_k] \). For the outputs, we write \( upv \) for vectors of zero-crossing expression variables \( [u_{p1}, \ldots, u_{pn}] \), \( xv \) for vectors of continuous state variables \( [x_{t1}, \ldots, x_{tk}] \) and \( dxv \) for the vectors of instantaneous derivatives \( [dx_{t1}, \ldots, dx_{tk}] \). The concatenation of two vectors is written \( [z_1, \ldots, z_n] @ [z'_1, \ldots, z'_m] \), it yields a new vector \( [z_1, z_1, \ldots, z_n, z'_1, \ldots, z'_m] \).

The translation is defined by \( \text{TrDef}(d) \), \( \text{TrEq}(E) \), \( \text{Tr}(e) \) and \( \text{TrAuto}(sp_i \rightarrow \ldots) \). We describe the first three, that treat data-flow primitives, before discussing the last.

The \( \text{TrDef}(d) \) function addresses function declarations, it relies on an auxiliary mapping \( \text{KindOf}(f) \) that gives the kind \( k \in \{A,D,C\} \) for the definition of a function \( f \). Discrete functions are not modified by the translation. Hybrid functions are altered by the addition of inputs and outputs which are passed through internal occurrences of zero-crossing operators \( up(e) \), ODEs \( \text{der} x = e \) and \( \text{init} x = e_0 \), and nested hybrid function applications \( f(e) \).

The \( \text{TrEq}(E) \) function treats equations. It returns a tuple \( (v, upv, lxv, xv, dxv, E') \), where the vectors contain the required inputs and outputs and \( E' \) is the translated equation (multiple equations are combined with \( \text{and} \)).

The \( \text{Tr}(e) \) function defines the translation of an expression. It returns a tuple \( (e', xv, upv, lxv, xv, dxv, E) \), with the usual vectors of inputs and outputs, and where \( e' \) is the translated expression and \( E \) are the auxiliary definitions.

These three functions are defined in Figure 4. To simplify the presentation, we assume non-conflicting names for all variables, and furthermore that for every variable \( x \) there are corresponding (and unique) variables: \( lx, last_x \) and \( dx \).

1. Variables \( y \) and constants \( v \) are not changed. They return empty vectors for all variables and equations.
2. Applications of \( \text{last}(y) \) are replaced with the corresponding initialized last value \( last_y \) whose eventual definition is guaranteed by the type system.
3. The operator \( up(e) \) is replaced by a fresh variable \( z \) that is added as a zero-crossing input. Another fresh variable \( u \) is defined by a new equation \( u = (e, \text{tt}) \) and added to the list of zero-crossing outputs, the second element of the pair is a status flag used in the translation of automata.\(^{11}\)
4. \( e_1 \text{ on } e_2 \) is translated into the boolean operation \&.
5. Local definitions \( \text{let } E \text{ in } e \) are flattened. This is sound, given appropriate renaming, in the basic language. But the compiler contains a more intricate treatment to avoid problems across automaton states, and to preserve the order of side-effects.
6. Applications of combinatorial (A) and discrete (D) functions are not changed.
7. For applications of hybrid (C) functions, we must account for the additional inputs and outputs added by the transformation. We thus introduce fresh input variables \( z \) and \( lx \), adding them to the appropriate vectors, a local variable \( r \) to hold the result of the function call, and three fresh output variables \( up, x \) and \( dx \). The local and output variables are defined by a new equation where the function is invoked, and the original expression is replaced by the variable holding the result. Note that the structure of nested hybrid function applications is reflected in the tree structure of the types of the new inputs and outputs.

\(^{11}\)Signals [4] are used instead of explicit status values in the full language where the added equation becomes \( \text{emit } u = e \).
\[ \text{Tr}(y) = \langle y, [\cdot], [\cdot], [\cdot], [\cdot], [\cdot], [\cdot] \rangle \]
\[ \text{Tr}(u) = \langle u, [\cdot], [\cdot], [\cdot], [\cdot], [\cdot], [\cdot] \rangle \]
\[ \text{Tr}(\text{last}(y)) = \langle \text{last}_y, [\cdot], [\cdot], [\cdot], [\cdot], [\cdot], [\cdot], [\cdot] \rangle \]
where \( \text{last}_y \) is the initialized last value corresponding to \( y \)
\[ \text{Tr}(\text{up}(e)) = \text{let } (e', z, u, \text{upv}, lx, xv, dxv, E) = \text{Tr}(e) \text{ in} \]
\[ (z, z, u, \text{upv}, lx, xv, dxv, u = (e', tt) \text{ and } E) \]
where \( z \) and \( u \) are fresh variables.
\[ \text{Tr}(e_1 \text{ on } e_2) = \text{let } (e'_1, zv_1, \text{upv}_1, lxv_1, xv_1, dxv_1, E_1) = \text{Tr}(e_1) \text{ in} \]
\[ (e'_1 \land e'_2, zv_2, \text{upv}_2, lxv_2, xv_2, dxv_2, E_2) = \text{Tr}(e_2) \text{ in} \]
\[ (e'_1 \land e'_2, zv_1 @ zv_2, \text{upv}_1 @ \text{upv}_2, lxv_1 @ lxv_2, xv_1 @ xv_2, dxv_1 @ dxv_2, E_1 \land E_2) \]
\[ \text{Tr}(\text{let } E \text{ in } e) = \text{let } (e', zv_1, \text{upv}_1, lxv_1, xv_1, dxv_1, E_1) = \text{TrEq}(E) \text{ in} \]
\[ (e', zv_2, \text{upv}_2, lxv_2, xv_2, dxv_2, E_2) = \text{Tr}(e) \text{ in} \]
\[ (e', zv_1 @ zv_2, \text{upv}_1 @ \text{upv}_2, lxv_1 @ lxv_2, xv_1 @ xv_2, dxv_1 @ dxv_2, E_1 \land E_2) \]
\[ \text{Tr}(f(e)) = \text{let } (f'(e'), zv, \text{upv}, lx, xv, dxv, E) = \text{Tr}(e) \text{ in} \]
\[ (f'(e'), zv, \text{upv}, lx, xv, dxv, E) \]
if \( \text{KindOf}(f) \in \{A, B\} \)
\[ \text{Tr}(f(e)) = \text{let } (e', zv, \text{upv}, lx, xv, dx, E) = \text{Tr}(e) \text{ in} \]
\[ (z, z, u, \text{upv}, lx, xv, dx, (r, up, x, dx) = f(z, lx, e') \text{ and } E) \]
if \( \text{KindOf}(f) = C \) and where \( r, z, u, lx, x, \) and \( dx \) are fresh variables.
\[ \text{TrEq}(x = e) = \text{let } (e', zv, \text{upv}, lx, xv, dxv, E) = \text{Tr}(e) \text{ in} \]
\[ (zv, \text{upv}, lx, xv, dxv, (x = e' \text{ and } E) \]
\[ \text{TrEq}(\text{init } x = e_0) = \text{let } (e'_0, zv_0, \text{upv}_0, lx_0, xv_0, dxv_0, E_0) = \text{Tr}(e_0) \text{ in} \]
\[ (zv_0, \text{upv}_0, lx_0, xv_0, dxv_0, (\text{last}_x = e'_0 \rightarrow lx) \text{ and } E_0) \]
where \( lx \) and \( \text{last}_x \) correspond with \( x \).
\[ \text{TrEq}(\text{der } x = e) = \text{let } (e', zv, \text{upv}, lx, xv, dx, E) = \text{Tr}(e) \text{ in} \]
\[ (zv, \text{upv}, lx, xv, dx, (x = \text{last}_x) \text{ and } (dx = e') \text{ and } E) \]
where \( lx, \text{last}_x \) and \( dx \) correspond with \( x \).
\[ \text{TrEq}(\text{automaton } (sp_i \longrightarrow \ldots)_{i \in I}) = \text{let } ((sp'_i \longrightarrow \ldots)_{i \in I}, zv, \text{upv}, lx, xv, dx, E) = \text{TrAuto}(\text{sp_i} \longrightarrow \ldots)_{i \in I} \text{ in} \]
\[ (zv, \text{upv}, lx, xv, dx, \text{automaton } (sp'_i \longrightarrow \ldots)_{i \in I}) \]
\[ \text{TrEq}(E_1 \text{ and } E_2) = \text{let } (zv_1, \text{upv}_1, lxv_1, xv_1, dxv_1, E'_1) = \text{TrEq}(E_1) \text{ in} \]
\[ let (zv_2, \text{upv}_2, lxv_2, xv_2, dxv_2, E'_2) = \text{TrEq}(E_2) \text{ in} \]
\[ (zv_1 @ zv_2, \text{upv}_1 @ \text{upv}_2, lxv_1 @ lxv_2, xv_1 @ xv_2, dxv_1 @ dxv_2, E'_1 \land E'_2) \]
\[ \text{TrDef}(\text{let } k \ f(y) = e) = \text{let } D f(xv, lxv, y) = \text{let } E \text{ in } (e', \text{upv}, xv, dxv) \text{ if } k \in \{A, B\} \]
\[ \text{TrDef}(\text{let } k \ f(y) = e) = \text{let } k \ f(y) = e \text{ if } k \in \{A, B\} \]
\[ \text{Translation of Automata.} \]

Automata are translated by a function
\[
\text{TrAuto}(\text{sp_i} \longrightarrow \text{u unless } (s_j)_{j \in I_i})_{i \in I}
\]
that returns a tuple
\[
((\text{sp}_i \longrightarrow \text{u' unless } (s'_j)_{j \in I_i})_{i \in I}, zv, lxv, \text{upv}, xv, dxv)
\]
The translation is only performed for automata in continuous contexts; discrete automata are not translated.

Some auxiliary definitions are used in the translation of automata. We write \( x \in v \) to mean that \( x \) is an element in
the vector \(u\), \(x\) and \(v_1 \notin v_2\) to mean that \(x \in v_1 \land x \notin v_2\). The union of vectors \(l_1\) and \(l_2\) is written \(l_1 \oplus l_2\), and defined:

\[
\begin{align*}
[l] \oplus l & = l \\
(x \oplus y) & = x \oplus (x \oplus y) \text{ if } x \notin y \\
(x \oplus y) & = x \oplus y \text{ otherwise}
\end{align*}
\]

This operator is lifted to sets of vectors:

\[
lx = \oplus_{i \in I} (lx_i) = lx_1 \oplus ... \oplus lx_n \text{ if } I = \{i_1, ..., i_n\}
\]

For a body \(u\) and a set of equations \(E\), \(\text{add}(u)(E)\) builds a new body from the given one by adding every \(e \in E\).

The translation of automata is presented in Figure 5. \(\text{TrAuto}(\cdot)\), translates the bodies and weak preemptions of all states \(i \in I\), via the auxiliary function \(\text{TrUntil}(\cdot)\). The vectors \(xv, upv, lxv, xv\) and \(dxv\) for an automaton are formed from the union of the corresponding vectors from each state body. All of the variables in the output vectors—\(upv, lxv, xv\) and \(dxv\)—have to be defined in every state body. For every variable \(up\) that is not already defined in a state, we add the equation \(up = (0.0, \text{ff})\); the reason is explained below. For an undefined derivative variable, we add the equation \(dxv = 0.0\). And for an undefined state variable, we add the equation \(x = lx\), where \(lx\) is the corresponding (uninitialized) state input. The strong transition expressions \((s_j)_{j \in J}\) are left unchanged since the guards are combinatorial and the actions are discrete.

\(\text{TrUntil}(\cdot)\) transforms the equations in body \(E\) but leaves the transition expressions \((s_j)_{j \in J}\) unchanged—we make the usual assumption that each weak transition guard is replaced with a variable defined in the state body. Local variable declarations are left in place, unless they apply to a continuous state variable, since these variables must be completed in the other states and available as outputs. But if \(x\) is revealed, the corresponding \(last_x\) variable must be hidden since otherwise it must also be defined in the other states.

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6. RELATED WORK AND DISCUSSION

Related Work.

There are other hybrid data-flow languages that include automata [3]. The two which have most influenced our approach are HYVISUAL/PTOLEMY II and Simulink/Stateflow. Lee and Zheng [10] have already considered hybrid systems models from a programming language perspective, and there are many similarities between their approach and ours. In particular, their implementation of HYVISUAL [13] on top of PTOLEMY II [14] allows arbitrary hierarchical nestings of automata, continuous data-flow networks, and discrete data-flow networks. Models are executed by cycling between discrete and continuous phases. A discrete phase is executed as long as new zero-crossings are detected. A continuous phase where the program is exercised by a numeric solver.

Changes to the execution model.

In the work presented in this paper, we try to extend our earlier approach [1] as conservatively as possible. Nevertheless, two adjustments to the original execution model are required when automata are added.

A hybrid program is executed by alternating between a continuous phase where the program is exercised by a numeric solver and a discrete phase where the effects of zero-crossings are calculated. Discrete phases succeed one another for as long as new zero-crossings are detected. An extra discrete reaction is required whenever a weak transition is fired to ensure that initialization of the successor state occurs in a discrete reaction and not inside the numeric solver. In the current implementation an extra discrete reaction is always executed after no more zero-crossings have been detected (and it may itself generate new zero-crossings).

The value of a zero-crossing output \(up\) becomes \((0.0, \text{ff})\) when the parent state of the zero-crossing is inactive. The 0.0 value ensures that a zero-crossing is not generated immediately after returning to the active state. Setting the status flag to \(\text{ff}\) tells the discrete solver not to generate a (spurious) zero-crossing even if the previous value of \(up\) was negative. Note that zero-crossing detection is not reset by a self-looping then transition.

12http://ptolemy.eecs.berkeley.edu/ptolemyII/
derivatives of continuous variables may only be set, using a \texttt{*\_dot} notation, in \textit{during} actions

The behavior of zero-crossings in \textsc{Stateflow} is also interesting. Transition guards in continuous automata are arbitrary boolean expressions—using input events makes a chart ‘triggered’ rather than continuous. When zero-crossings are enabled for a chart, the simulation engine tries to accurately detect the instants of transition, not by monitoring the details of sub-expressions, but instead by changing the value of a zero-crossing expression from negative to positive upon state entry.\footnote{This is mentioned in the documentation; and the insensitivity to guard dynamics is also evident in debugger traces.} Charts are executed just before \((t - \epsilon)\) the calculated time of a zero-crossing \((t)\), and again just afterward \((t + \epsilon)\). In this way, cascades of zero-crossings are also possible, but some time always passes between any two.

While \textsc{Simulink/Stateflow} is obviously much richer and more polished than our proposal, we believe that our approach is not without value. For instance, rather than effectively having two simulation engines \cite[p. 28]{3}, we seek a consistent and simple integration of continuous automata and data-flow networks. Furthermore, our goal of better understanding the fundamentals of such tools mandates that we seek a small set of primitives, rather than attempt to provide a large number of features.

**Boolean guards.**

The encoding for boolean guards mentioned in Remark 1 is enough to detect enabling in a continuous phase, but not to have transitions with enabled guards taken continually. This alternative can be implemented using intricate encodings of zero-crossings, but it would also be possible to either change the definition of \texttt{up}(x) to \((x \geq 0 \rightarrow z)\), that is true after a reset if \(x\) is not negative and then equal to the input from the solver, or to add primitives from which the same behaviour can be constructed, i.e. to add an \texttt{init: bool} \rightarrow \texttt{zero}, that only tests its argument when reset, and an \texttt{orz: zero \times zero} \rightarrow \texttt{zero} for combining events.

**Compiling automata first.**

This paper presents one of the paths shown in Figure 1, but we have also experimented with translating automata before ODEs. Our type system is not fine enough, however, to maintain the information implicit to an automaton that guarantees sound interaction between discrete and continuous behaviors. In simple terms: the intermediate results of compilation are not well-typed. This is a subject of our continuing research.

**Sharing solver resources.**

Another subject of our continuing research is to find a natural way of sharing zero-crossings and continuous states across mutually-exclusive automaton modes. This turns out to be quite delicate in our current execution and semantic models, and it is an inadequacy that we would like to correct.

7. CONCLUSION

This paper introduced a language for programming explicit hybrid systems with synchronous concurrency, ordinary differential equations, and hierarchical automata. A synchronous language is used at the programming level and as a target for code generation via a source-to-source translation. Unsound combinations of continuous and discrete elements are rejected by a type system. Our approach elucidates some aspects of designing hybrid modelers.

8. REFERENCES

\begin{itemize}
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