Real-Time Systems

Lecture 9: DC Properties IIa

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Contents & Goals

Last Lecture:

- DC Implementables

This Lecture:

- **Educational Objectives:** Capabilities for following tasks/questions.
  - Facts: (un)decidability properties of DC in discrete/continuous time.
  - What's the idea of the considered (un)decidability proofs?

- **Content:**
  - RDC in discrete time cont'd
  - Satisfiability and realisability from 0 is decidable for RDC in discrete time
  - Undecidable problems of DC in continuous time

RDC in Discrete Time Cont'd

Restricted DC (RDC)

$$F ::= \lceil P \rceil \mid \neg F_1 \mid F_1 \vee F_2 \mid F_1 \,;\, F_2$$

where P is a state assertion, but with **boolean** observables **only**.

Note:

- No global variables, thus don't need $\mathcal{V}$.
-

Discrete Time Interpretations

- An interpretation $\mathcal{I}$ is called **discrete time interpretation** if and only if, for each state variable X,

$$X_{\mathcal{I}} : \mathsf{Time} \rightarrow \mathcal{D}(X)$$

  with

  - $\mathsf{Time} = \mathbb{R}_0^+$,
  - all discontinuities are in $\mathbb{N}_0$.

- An interval $[b, e] \subset \mathsf{Intv}$ is called **discrete** if and only if $b, e \in \mathbb{N}_0$.
- We say (for a discrete time interpretation $\mathcal{I}$ and a discrete interval $[b, e]$)

$$\mathcal{I}, [b, e] \models F_1 \,;\, F_2$$

  if and only if there exists $m \in [b, e] \cap \mathbb{N}_0$ such that

$$\mathcal{I}, [b, m] \models F_1 \qquad \text{and} \qquad \mathcal{I}, [m, e] \models F_2$$

Differences between Continuous and Discrete Time

- Let P be a state assertion.

	Continuous Time	Discrete Time
$\models^? (\lceil P \rceil \,;\, \lceil P \rceil) \implies \lceil P \rceil$	✔	✔
$\models^? \lceil P \rceil \implies (\lceil P \rceil \,;\, \lceil P \rceil)$	✔	✘

- In particular: $\ell = 1 \iff (\lceil 1 \rceil \wedge \neg(\lceil 1 \rceil \,;\, \lceil 1 \rceil))$ (in discrete time).

Expressiveness of RDC

- $\ell = 1 \iff \lceil 1 \rceil \wedge \neg(\lceil 1 \rceil \,;\, \lceil 1 \rceil)$
- $\ell = 0 \iff \neg\lceil 1 \rceil$
- $true \iff \ell = 0 \vee \neg(\ell = 0)$
- $\int P = 0 \iff \lceil \neg P \rceil \vee \ell = 0$
- $\int P = 1 \iff (\int P = 0) \,;\, (\lceil P \rceil \wedge \ell = 1) \,;\, (\int P = 0)$
- $\int P = k + 1 \iff (\int P = k) \,;\, (\int P = 1)$
- $\int P \geq k \iff (\int P = k) \,;\, true$
- $\int P > k \iff \int P \geq k + 1$
- $\int P \leq k \iff \neg(\int P > k)$
- $\int P < k \iff \int P \leq k - 1$

where $k \in \mathbb{N}$.

Decidability of Satisfiability/Realisability from 0

Theorem 3.6.
The satisfiability problem for RDC with discrete time is decidable.

Theorem 3.9.
The realisability problem for RDC with discrete time is decidable.

$\exists\, \mathcal{I}, [b,e] \bullet \mathcal{I}, [b,e] \models F$?

construct

$\mathcal{L}(F)$ regular

procedure

$\mathcal{L}(F) \neq \emptyset$: jedes $w \in \mathcal{L}(F)$ definiert ein Bsp. für $\mathcal{I}, [b,e] \models F$

$\mathcal{L}(F) = \emptyset$: F not satisfiable

$\mathcal{L}(F) \overset{\text{bij}}{\longleftrightarrow} \{\mathcal{I} \mid \mathcal{I} \models F\}$

decidable

Sketch: Proof of Theorem 3.6

- give a procedure to construct, given a formula F, a **regular** language $\mathcal{L}(F)$ such that

$$\mathcal{I}, [0, n] \models F \text{ if and only if } w \in \mathcal{L}(F)$$

  where word w describes $\mathcal{I}$ on $[0, n]$
  (suitability of the procedure: **Lemma 3.4**)

- then F is satisfiable in discrete time if and only if $\mathcal{L}(F)$ is not empty (**Lemma 3.5**)

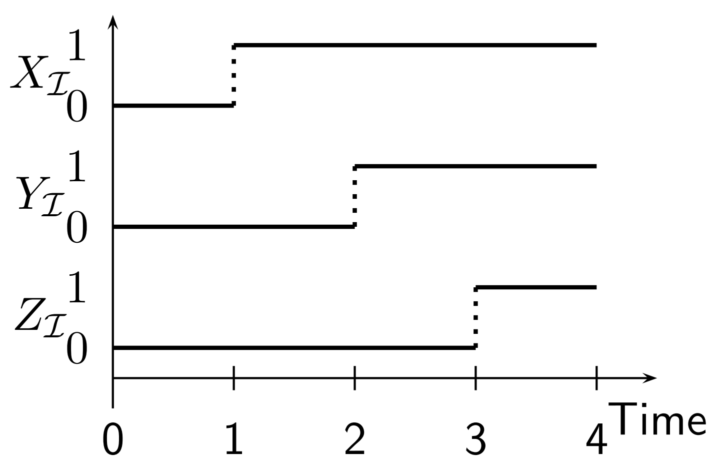
- Theorem 3.6 follows because
  - $\mathcal{L}(F)$ can **effectively** be constructed,
  - the emptyness problem is **decidable** for regular languages.

Construction of $\mathcal{L}(F)$

- **Idea:**
  - alphabet $\Sigma(F)$ consists of basic conjuncts of the state variables in F,
  - a letter corresponds to an interpretation on an interval of length 1,
  - a word of length n describes an interpretation on interval $[0, n]$.

- **Example:** Assume F contains exactly state variables X, Y, Z, then

$$\Sigma(F) = \{X \wedge Y \wedge Z, X \wedge Y \wedge \neg Z, X \wedge \neg Y \wedge Z, X \wedge \neg Y \wedge \neg Z, \neg X \wedge Y \wedge Z, \neg X \wedge Y \wedge \neg Z, \neg X \wedge \neg Y \wedge Z, \neg X \wedge \neg Y \wedge \neg Z\}.$$



$$w = (\neg X \wedge \neg Y \wedge \neg Z) \cdot (X \wedge \neg Y \wedge \neg Z) \cdot (X \wedge Y \wedge \neg Z) \cdot (X \wedge Y \wedge Z) \in \Sigma(F)^*$$

Concatenation

Construction of $\mathcal{L}(F)$ more Formally

Definition 3.2. A word $w = a_1 \ldots a_n \in \Sigma(F)^*$ with $n \geq 0$ **describes** a **discrete** interpretation $\mathcal{I}$ on $[0, n]$ if and only if

$$\forall j \in \{1, \ldots, n\}\ \forall t \in\ ]j-1, j[\ : \mathcal{I}[\![a_j]\!](t) = 1.$$

For $n = 0$ we put $w = \varepsilon$.

$P = X \wedge \neg Y \iff (X \wedge \neg Y \wedge Z) \vee (X \wedge \neg Y \wedge \neg Z)$

- Each state assertion P can be transformed into an equivalent **disjunctive normal form** $\bigvee_{i=1}^{m} a_i$ with $a_i \in \Sigma(F)$.
- Set $DNF(P) := \{a_1, \ldots, a_m\}$ $(\subseteq \Sigma(F))$.
- Define $\mathcal{L}(F)$ inductively:

$DNF(X \wedge \neg Y) = \{X \wedge \neg Y \wedge Z, X \wedge \neg Y \wedge \neg Z\}$

finite words, length at least one

$$\mathcal{L}(\lceil P \rceil) = DNF(P)^+,$$
$$\mathcal{L}(\neg F_1) = \Sigma(F)^* \setminus \mathcal{L}(F_1),$$
$$\mathcal{L}(F_1 \vee F_2) = \mathcal{L}(F_1) \cup \mathcal{L}(F_2),$$
$$\mathcal{L}(F_1 \,;\, F_2) = \mathcal{L}(F_1) \cdot \mathcal{L}(F_2).$$

$\leadsto \mathcal{L}(F)$ regular

Lemma 3.4

Lemma 3.4. For all RDC formulae F, discrete interpretations $\mathcal{I}$, $n \geq 0$, and all words $w \in \Sigma(F)^*$ which **describe** $\mathcal{I}$ on $[0, n]$,

$$\mathcal{I}, [0, n] \models F \text{ if and only if } w \in \mathcal{L}(F).$$

Proof: Structural induction.

Base $F = \lceil P \rceil$: Let $w = a_1, \dots, a_n$, $n \geq 0$, describe $\mathcal{I}$ on $[0,n]$.

$\mathcal{I}, [0,n] \models \lceil P \rceil \iff \mathcal{I}, [0,n] \models \lceil P \rceil$ and $n \geq 1$

$\iff n \geq 1$ and $\forall\, 1 \leq j \leq n \bullet \mathcal{I}, [j-1, j] \models \lceil P \rceil$

$\iff n \geq 1$ and $\forall\, 1 \leq j \leq n \bullet \mathcal{I}, [j-1, j] \models \lceil P \rceil \wedge \lceil a_j \rceil$ and $a_j \in DNF(P)$ (describes)

$\iff n \geq 1$ and $\forall\, 1 \leq j \leq n \bullet a_j \in DNF(P)$ (clear)

$\iff w \in DNF(P)^+$

$\iff w \in \mathcal{L}(\lceil P \rceil)$

Steps:
- $\neg F_1$
- $F_1 \vee F_2$
- $F_1 \,;\, F_2$

Sketch: Proof of Theorem 3.9

Theorem 3.9.
The realisability problem for RDC with discrete time is decidable.

- $kern(L)$ contains all words of L whose prefixes are again in L.
- If L is regular, then $kern(L)$ is also regular.
- $kern(\mathcal{L}(F))$ can effectively be constructed.
- We have

Lemma 3.8. For all RDC formulae F, F is realisable from 0 in discrete time if and only if $kern(\mathcal{L}(F))$ is infinite.

- Infinity of regular languages is decidable.

(Variants of) RDC in Continuous Time

Recall: Restricted DC (RDC)

$$F ::= \lceil P \rceil \mid \neg F_1 \mid F_1 \vee F_2 \mid F_1 \,;\, F_2$$

where P is a state assertion, but with **boolean** observables **only**.

From now on: "RDC $+ \ell = x, \forall\, x$"

$$F ::= \lceil P \rceil \mid \neg F_1 \mid F_1 \vee F_2 \mid F_1 \,;\, F_2 \mid \ell = 1 \mid \ell = x \mid \forall\, x \bullet F_1$$

Undecidability of Satisfiability/Realisability from 0

Theorem 3.10.
The realisability from 0 problem for DC with **continuous time** is undecidable, not even semi-decidable.

Theorem 3.11.
The satisfiability problem for DC with continuous time is undecidable.

Sketch: Proof of Theorem 3.10

Reduce divergence of **two-counter machines** to realisability from 0:

- Given a two-counter machine $\mathcal{M}$ with final state q_{fin},
- construct a DC formula $F(\mathcal{M}) := encoding(\mathcal{M})$
- such that

$\mathcal{M}$ **diverges** **if and only if** the DC formula

$$F(\mathcal{M}) \wedge \neg\Diamond\lceil q_{fin}\rceil$$

is **realisable from 0**.

- If realisability from 0 was (semi-)decidable,
  divergence of two-counter machines would be (which it isn't).

Recall: Two-counter machines

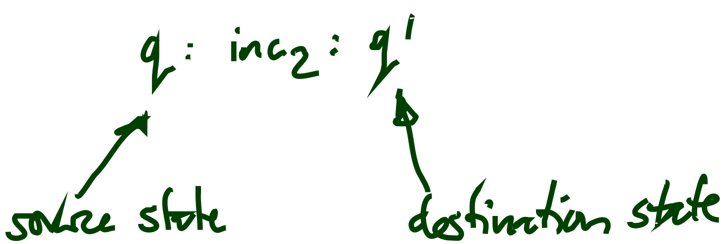
A **two-counter** machine is a structure

$$\mathcal{M} = (\mathcal{Q}, q_0, q_{fin}, Prog)$$

where

- $\mathcal{Q}$ is a finite set of **states**,
- comprising the **initial state** q_0 and the **final state** q_{fin}
- $Prog$ is the **machine program**, i.e. a finite set of **commands** of the form

$$q : inc_i : q' \quad \text{and} \quad q : dec_i : q', q'', \qquad i \in \{1, 2\}.$$



- We assume **deterministic** 2CM: for each $q \in \mathcal{Q}$, at most one command starts in q, and q_{fin} is the only state where no command starts.

2CM Configurations and Computations

current state → values of counters

- a **configuration** of $\mathcal{M}$ is a triple $K = (q, n_1, n_2) \in \mathcal{Q} \times \mathbb{N}_0 \times \mathbb{N}_0$.
- The **transition relation** "$\vdash$" on configurations is defined as follows:

Command	Semantics: $K \vdash K'$
$q : inc_1 : q'$	$(q, n_1, n_2) \vdash (q', n_1 + 1, n_2)$
$q : dec_1 : q', q''$	$(q, 0, n_2) \vdash (q', 0, n_2)$
	$(q, n_1 + 1, n_2) \vdash (q'', n_1, n_2)$
$q : inc_2 : q'$	$(q, n_1, n_2) \vdash (q', n_1, n_2 + 1)$
$q : dec_2 : q', q''$	$(q, n_1, 0) \vdash (q', n_1, 0)$
	$(q, n_1, n_2 + 1) \vdash (q'', n_1, n_2)$

- The (!) **computation** of $\mathcal{M}$ is a finite sequence of the form ("$\mathcal{M}$ **halts**")

$$K_0 = (q_0, 0, 0) \vdash K_1 \vdash K_2 \vdash \cdots \vdash (q_{fin}, n_1, n_2)$$

or an infinite sequence of the form ("$\mathcal{M}$ **diverges**")

$$K_0 = (q_0, 0, 0) \vdash K_1 \vdash K_2 \vdash \ldots$$

2CM Example

- $\mathcal{M} = (\mathcal{Q}, q_0, q_{fin}, \mathit{Prog})$
- commands of the form $q : inc_i : q'$ and $q : dec_i : q', q''$, $i \in \{1, 2\}$
- configuration $K = (q, n_1, n_2) \in \mathcal{Q} \times \mathbb{N}_0 \times \mathbb{N}_0$.
-

Command	Semantics: $K \vdash K'$
$q : inc_1 : q'$	$(q, n_1, n_2) \vdash (q', n_1 + 1, n_2)$
$q : dec_1 : q', q''$	$(q, 0, n_2) \vdash (q', 0, n_2)$ $(q, n_1 + 1, n_2) \vdash (q'', n_1, n_2)$
$q : inc_2 : q'$	$(q, n_1, n_2) \vdash (q', n_1, n_2 + 1)$
$q : dec_2 : q', q''$	$(q, n_1, 0) \vdash (q', n_1, 0)$ $(q, n_1, n_2 + 1) \vdash (q'', n_1, n_2)$

$Q = \{q_0, q_1, q_{fin}\}$

$Prog = \{q_0 : inc_1 : q_1,$
$q_1 : inc_1 : q_{fin}\}$

$(q_0, 0, 0)$
$\vdash$
$(q_1, 1, 0)$
$\vdash$
$(q_{fin}, 2, 0)$

→ machine halts

$Q = \{q_0, q_{fin}\}$

$Prog = \{q_0, inc_2, q_0\}$

$(q_0, 0, 0)$
$\vdash$
$(q_0, 0, 1)$
$\vdash$
$(q_0, 0, 2)$
⋮

→ machine diverges

Reducing Divergence to DC realisability: Idea In Pictures

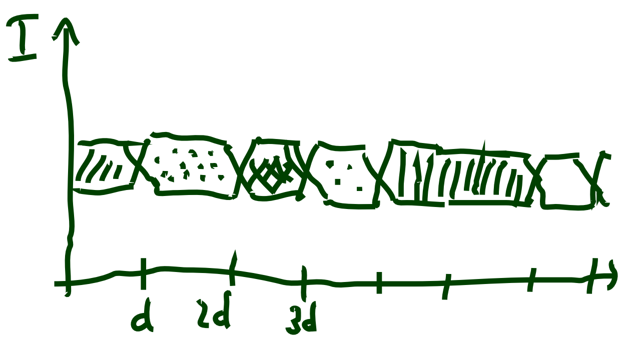
2CM $\mathcal{M}$ diverges

iff ✓

exists $\pi : k_0 \vdash k_1 \vdash \ldots$

iff:

exist



("I describes π")

and

$\mathcal{I} \models_0 \mathcal{F}(\mathcal{M}) \wedge \neg \Diamond \lceil q_{fin} \rceil$

$\mathcal{F}(\mathcal{M})$ intuitively requires:

- $[0, d]$ encodes $(q_0, 0, 0)$
- $[n \cdot d, (n+1) \cdot d]$ encodes a configuration
- $[n \cdot d, (n+1) \cdot d]$ and $[(n+1) \cdot d, (n+2) \cdot d]$ encode configurations which are in $\vdash$-Relation
- if q_{fin} is reached, we stay there

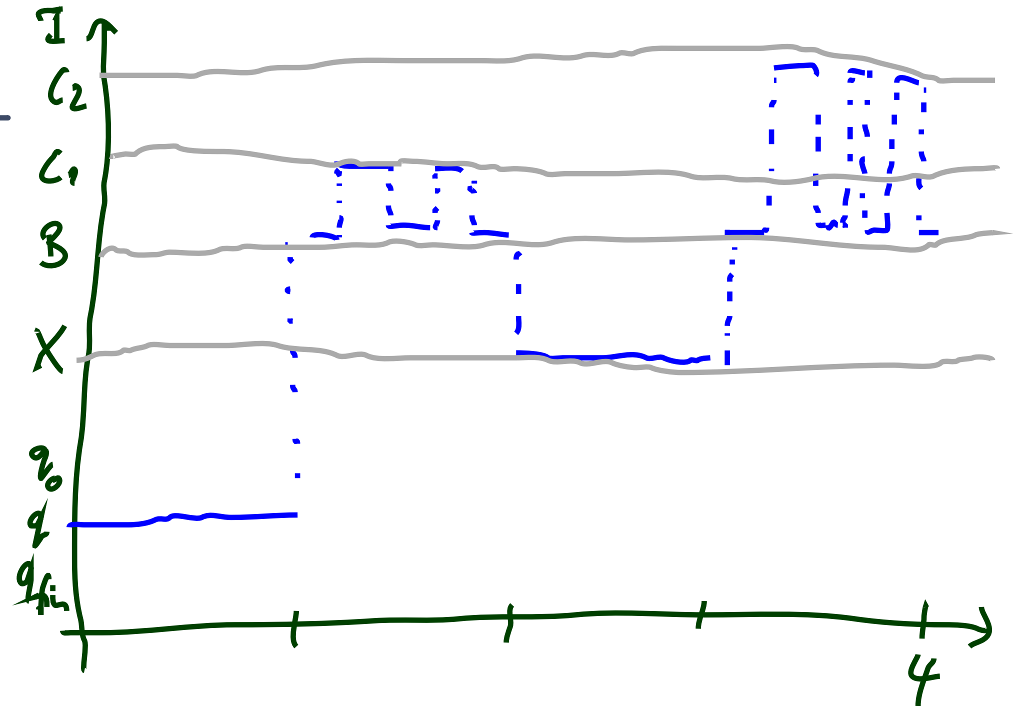
Reducing Divergence to DC realisability: Idea

- A single configuration K of $\mathcal{M}$ can be encoded in an interval of length 4; being an encoding interval can be **characterised** by a DC formula.
- An interpretation on 'Time' encodes **the** computation of $\mathcal{M}$ if
  - each interval $[4n, 4(n+1)]$, $n \in \mathbb{N}_0$, **encodes** a configuration K_n,
  - each two subsequent intervals $[4n, 4(n+1)]$ and $[4(n+1), 4(n+2)]$, $n \in \mathbb{N}_0$, encode configurations $K_n \vdash K_{n+1}$ **in transition relation**.
- Being encoding of the run can be **characterised** by DC formula $F(\mathcal{M})$.
- Then $\mathcal{M}$ **diverges** if and only if $F(\mathcal{M}) \wedge \neg\Diamond\lceil q_{\mathit{fin}} \rceil$ is realisable from 0.

Encoding Configurations

- We use Obs = {obs} with
  $\mathcal{D}(\text{obs}) = \mathcal{Q}_{\mathcal{M}} \,\dot{\cup}\, \{C_1, C_2, B, X\}$.

states of $\mathcal{M}$ — *disjoint union*



Examples:

- $K = (q, 2, 3)$

$$\left(\begin{array}{c}\lceil q \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil B \rceil ; \lceil C_1 \rceil ; \lceil B \rceil ; \lceil C_1 \rceil ; \lceil B \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil X \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil B \rceil ; \lceil C_2 \rceil ; \lceil B \rceil ; \lceil C_2 \rceil ; \lceil B \rceil ; \lceil C_2 \rceil ; \lceil B \rceil\\ \wedge\\ \ell = 1\end{array}\right)$$

- $K_0 = (q_0, 0, 0)$

$$\left(\begin{array}{c}\lceil q_0 \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil B \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil X \rceil\\ \wedge\\ \ell = 1\end{array}\right) ; \left(\begin{array}{c}\lceil B \rceil\\ \wedge\\ \ell = 1\end{array}\right)$$

or, using abbreviations, $\lceil q_0 \rceil^1 ; \lceil B \rceil^1 ; \lceil X \rceil^1 ; \lceil B \rceil^1$.

Construction of $F(\mathcal{M})$

In the following, we give DC formulae describing

- the initial configuration,
- the general form of configurations,
- the transitions between configurations,
- the handling of the final state.

$F(\mathcal{M})$ is the conjunction of all these formulae.

Initial and General Configurations

$$init :\Longleftrightarrow (\ell \geq 4 \implies \lceil q_0 \rceil^1 \,;\, \lceil B \rceil^1 \,;\, \lceil X \rceil^1 \,;\, \lceil B \rceil^1 \,;\, true)$$

$$\begin{aligned} keep :\Longleftrightarrow \Box(\lceil Q \rceil^1 \,;\, \lceil B \vee C_1 \rceil^1 \,;\, \lceil X \rceil^1 \,;\, \lceil B \vee C_2 \rceil^1 \,;\, \ell = 4 \\ \implies \ell = 4 \,;\, \lceil Q \rceil^1 \,;\, \lceil B \vee C_1 \rceil^1 \,;\, \lceil X \rceil^1 \,;\, \lceil B \vee C_2 \rceil^1) \end{aligned}$$

where $Q := \neg(X \vee C_1 \vee C_2 \vee B)$.

Auxiliary Formula Pattern $copy$

$$copy(F, \{P_1, \dots, P_n\}) :\iff$$

$$\forall\, c, d \bullet \Box((F \wedge \ell = c) \,;\, (\lceil P_1 \vee \dots \vee P_n \rceil \wedge \ell = d) \,;\, \lceil P_1 \rceil \,;\, \ell = 4$$
$$\implies \ell = c + d + 4 \,;\, \lceil P_1 \rceil$$

$$\dots$$

$$\forall\, c, d \bullet \Box((F \wedge \ell = c) \,;\, (\lceil P_1 \vee \dots \vee P_n \rceil \wedge \ell = d) \,;\, \lceil P_n \rceil \,;\, \ell = 4$$
$$\implies \ell = c + d + 4 \,;\, \lceil P_n \rceil$$

$q : inc_1 : q'$ (Increment)

(i) Change state

$$\Box(\lceil q \rceil^1 \,;\, \lceil B \vee C_1 \rceil^1 \,;\, \lceil X \rceil^1 \,;\, \lceil B \vee C_2 \rceil^1 \,;\, \ell = 4 \implies \ell = 4 \,;\, \lceil q' \rceil^1 \,;\, true)$$

(ii) Increment counter

$$\forall\, d \bullet \Box(\lceil q \rceil^1 \,;\, \lceil B \rceil^d \,;\, (\ell = 0 \vee \lceil C_1 \rceil \,;\, \lceil \neg X \rceil) \,;\, \lceil X \rceil^1 \,;\, \lceil B \vee C_2 \rceil^1 \,;\, \ell = 4$$
$$\implies \ell = 4 \,;\, \lceil q' \rceil^1 \,;\, (\lceil B \rceil \,;\, \lceil C_1 \rceil \,;\, \lceil B \rceil \wedge \ell = d) \,;\, true$$

$q : inc_1 : q'$ (Increment)

(i) Keep rest of first counter

$$copy(\lceil q \rceil^1 \,;\, \lceil B \vee C_1 \rceil \,;\, \lceil C_1 \rceil, \{B, C_1\})$$

(ii) Leave second counter unchanged

$$copy(\lceil q \rceil^1 \,;\, \lceil B \vee C_1 \rceil \,;\, \lceil X \rceil^1, \{B, C_2\})$$

$q : dec_1 : q', q''$ *(Decrement)*

(i) If zero

$$\Box(\lceil q \rceil^1 ; \lceil B \rceil^1 ; \lceil X \rceil^1 ; \lceil B \vee C_2 \rceil^1 ; \ell = 4 \implies \ell = 4 ; \lceil q' \rceil^1 ; \lceil B \rceil^1 ; true)$$

(ii) Decrement counter

$$\forall\, d \bullet \Box(\lceil q \rceil^1 ; (\lceil B \rceil ; \lceil C_1 \rceil \wedge \ell = d) ; \lceil B \rceil ; \lceil B \vee C_1 \rceil ; \lceil X \rceil^1 ; \lceil B \vee C_2 \rceil^1 ; \ell = \\ \implies \ell = 4 ; \lceil q'' \rceil^1 ; \lceil B \rceil^d ; true)$$

(iii) Keep rest of first counter

$$copy(\lceil q \rceil^1 ; \lceil B \rceil ; \lceil C_1 \rceil ; \lceil B_1 \rceil, \{B, C_1\})$$

Final State

$$copy(\lceil q_{\mathit{fin}} \rceil^1 \,;\, \lceil B \vee C_1 \rceil^1 \,;\, \lceil X \rceil \,;\, \lceil B \vee C_2 \rceil^1, \{q_{\mathit{fin}}, B, X, C_1, C_2\})$$

Satisfiability

- Following [Chaochen and Hansen, 2004] we can observe that

  $\mathcal{M}$ **halts if and only if** the DC formula $F(\mathcal{M}) \land \Diamond\lceil q_{\mathit{fin}} \rceil$ is **satisfiable**.

  This yields

  > **Theorem 3.11.** The satisfiability problem for DC with continuous time is undecidable.

  (It is semi-decidable.)

- Furthermore, by taking the contraposition, we see

  $\mathcal{M}$ **diverges** **if and only if** $\mathcal{M}$ does not **halt**
  if and only if $F(\mathcal{M}) \land \neg\Diamond\lceil q_{\mathit{fin}} \rceil$ is **not** satisfiable.

- Thus whether a DC formula is **not satisfiable** is not decidable, not even semi-decidable.

Validity

- By Remark 2.13, F is valid iff $\neg F$ is not satisfiable, so

> **Corollary 3.12.** The validity problem for DC with continuous time is undecidable, not even semi-decidable.

Validity

- By Remark 2.13, F is valid iff $\neg F$ is not satisfiable, so

> **Corollary 3.12.** The validity problem for DC with continuous time is undecidable, not even semi-decidable.

- This provides us with an alternative proof of Theorem 2.23 ("there is no sound and complete proof system for DC"):

Validity

- By Remark 2.13, F is valid iff $\neg F$ is not satisfiable, so

> **Corollary 3.12.** The validity problem for DC with continuous time is undecidable, not even semi-decidable.

- This provides us with an alternative proof of Theorem 2.23 ("there is no sound and complete proof system for DC"):
  - **Suppose** there were such a calculus $\mathcal{C}$.
  - By Lemma 2.22 it is semi-decidable whether a given DC formula F is a theorem in $\mathcal{C}$.
  - By the soundness and completeness of $\mathcal{C}$, F is a theorem in $\mathcal{C}$ **if and only if** F is valid.
  - Thus it is semi-decidable whether F is valid. **Contradiction.**

Discussion

- Note: the DC fragment defined by the following grammar is **sufficient** for the reduction

$$F ::= \lceil P \rceil \mid \neg F_1 \mid F_1 \vee F_2 \mid F_1 \,;\, F_2 \mid \ell = 1 \mid \ell = x \mid \forall\, x \bullet F_1,$$

  P a state assertion, x a global variable.

- Formulae used in the reduction are abbreviations:

$$\begin{aligned} \ell = 4 &\iff \ell = 1 \,;\, \ell = 1 \,;\, \ell = 1 \,;\, \ell = 1 \\ \ell \geq 4 &\iff \ell = 4 \,;\, \mathit{true} \\ \ell = x + y + 4 &\iff \ell = x \,;\, \ell = y \,;\, \ell = 4 \end{aligned}$$

- Length 1 is not necessary — we can use $\ell = z$ instead, with fresh z.
- This is RDC augmented by "$\ell = x$" and "$\forall\, x$", which we denote by **RDC** $+ \ell = x, \forall\, x$.

References

[Chaochen and Hansen, 2004] Chaochen, Z. and Hansen, M. R. (2004). *Duration Calculus: A Formal Approach to Real-Time Systems*. Monographs in Theoretical Computer Science. Springer-Verlag. An EATCS Series.

[Olderog and Dierks, 2008] Olderog, E.-R. and Dierks, H. (2008). *Real-Time Systems - Formal Specification and Automatic Verification*. Cambridge University Press.