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Resilience of Timed Systems

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Abstract

This paper addresses reliability of timed systems in the setting of *resilience*, that considers the behaviors of a system when unspecified timing errors such as missed deadlines occur. Given a fault model that allows transitions to fire later than allowed by their guard, a system is *universally resilient* (or self-resilient) if after a fault, it always returns to a timed behavior of the non-faulty system. It is *existentially resilient* if after a fault, there exists a way to return to a timed behavior of the non-faulty system, that is, if there exists a controller which can guide the system back to a normal behavior. We show that universal resilience of timed automata is undecidable, while existential resilience is decidable, in EXPSPACE. To obtain better complexity bounds and decidability of universal resilience, we consider untimed resilience, as well as subclasses of timed automata.

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1 Introduction

Timed automata [2] are a natural model for cyber-physical systems with real-time constraints that have led to an enormous body of theoretical and practical work. Formally, timed automata are finite-state automata equipped with real valued variables called clocks, that measure time and can be reset. Transitions are guarded by logical assertions on the values of these clocks, which allows for the modeling of real-time constraints, such as the time elapsed between the occurrence of two events. A natural question is whether a real-time system can handle unexpected delays. This is a crucial need when modeling systems that must follow a priori schedules such as trains, metros, buses, etc. Timed automata are not a priori tailored to handle unspecified behaviors: guards are mandatory time constraints, i.e., transition firings must occur within the prescribed delays. Hence, transitions cannot occur late, except if late transitions are explicitly specified in the model. This paper considers the question of resilience for timed automata, i.e., study whether a system returns to its normal specified timed behavior after an unexpected but unavoidable delay.

Several works have addressed timing errors as a question of *robustness* [10, 8, 7], to guarantee that a property of a system is preserved for some small imprecision of up to ϵ time units. Timed automata have an ideal representation of time: if a guard of a transition contains a constraint of the form $x = 12$, it means that this transition occurs *exactly* when the value of clock x is 12. Such an arbitrary precision is impossible in an implementation [10]. One way of addressing this is through guard enlargement, i.e., by checking that there exists a small value $\epsilon > 0$ such that after replacing guards of the form $x \in [a, b]$ by $x \in [a - \epsilon, b + \epsilon]$,



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	Universal Resilience	Existential Resilience
Timed	Undecidable for TA (Prop. 18) EXPSPACE-C for IRTA (Thm. 20)	EXPSPACE (Thm. 14) PSPACE-Hard (Thm. 15, Thm. 32)
Untimed	EXPSPACE-C (Thm. 21)	PSPACE-C (Thm. 16, Rmk. 17)

■ **Table 1** Summary of results for resilience.

46 the considered property is still valid, as shown in [7] for ω -regular properties. In [15], robust
47 automata are defined that accept timed words and their neighbors i.e., words whose timing
48 differences remain at a small distance, while in [16, 12, 19, 1], the authors consider robustness
49 via modeling clock drifts. Our goal is different: rather than being robust w.r.t. to slight
50 imprecisions, we wish to check the capacity to recover from a *possibly large* time deviation.
51 Thus, for a bounded number of steps, the system can deviate arbitrarily, after which, it must
52 return to its specified timed behavior.

53 A first contribution of this paper is a formalization of resilience in timed automata. We
54 capture delayed events with *faulty transitions*. These occur at dates deviating from the
55 original specification and may affect clock values for an arbitrarily long time, letting the
56 system diverge from its expected behavior. A system is *resilient* if it recovers in a finite
57 number of steps after the fault. More precisely, we define two variants. A timed automaton
58 is *K - \forall -resilient* if for *every* faulty timed run, the behavior of the system K steps after the
59 fault cannot be distinguished from a non-faulty behavior. In other words, the system *always*
60 repairs itself in at most K steps after a fault, whenever a fault happens. This means that,
61 after a fault happens, *all* the subsequent behaviors (or extensions) of the system are restored
62 to normalcy within K steps. A timed automaton is *K - \exists -resilient* if for every timed run
63 ending with a fault, there exists *an* extension in which, the behavior of the system K steps
64 after the fault cannot be distinguished from a non-faulty behavior. There can still be some
65 extensions which are beyond repair, or take more than K steps after fault to be repaired,
66 but there is a guarantee of at least one repaired extension within K steps after the fault.
67 In the first case, the timed automaton is fully self-resilient, while in the second case, there
68 exist controllers choosing dates and transitions so that the system gets back to a normal
69 behavior. We also differentiate between timed and untimed settings: in timed resilience
70 recovered behaviors must be indistinguishable w.r.t. actions and dates, while in untimed
71 resilience recovered behaviors only need to match actions.

72 Our results are summarized in Table 1: we show that the question of universal resilience
73 and inclusion of timed languages are inter-reducible. Thus *timed* universal resilience is
74 undecidable in general, and decidable for classes for which inclusion of timed languages
75 is decidable and which are stable under our reduction. This includes the class of Integer
76 Reset Timed Automata (IRTA) [18] for which we obtain EXPSPACE containment. Further,
77 *untimed* universal resilience is EXPSPACE-Complete in general.

78 Our main result concerns existential resilience, which requires new non-trivial core
79 contributions because of the quantifier alternation ($\forall\exists$). The classical region construction
80 is not precise enough: we introduce *strong regions* and develop novel techniques based on
81 these, which ensure that all runs following a strong region have (i) matching integral time
82 elapses, and (ii) the fractional time can be retimed to visit the same set of locations and
83 (usual) regions. Using this technique, we show that existential timed resilience is decidable,
84 in EXPSPACE. We also show that untimed existential resilience is PSPACE-Complete.

85 **Related Work:** Resilience has been considered with different meanings: In [13], faults
86 are modeled as conflicts, the system and controller as *deterministic* timed automata, and
87 avoiding faults reduces to checking reachability. This is easier than universal resilience which
88 reduces to timed language inclusion, and existential resilience which requires a new notion of

89 regions. In [14] a system, modeled as an untimed I/O automaton, is considered “sane” if its
 90 runs contain at most k errors, and allow a sufficient number s of error-free steps between two
 91 violations of an LTL property. It is shown how to synthesize a sane system, and compute
 92 (Pareto-optimal) values for s and k . In [17], the objective is to synthesize a transducer E ,
 93 possibly with memory, that reads a timed word σ produced by a timed automaton \mathcal{A} , and
 94 outputs a timed word $E(\sigma)$ obtained by deleting, delaying or forging new timed events, such
 95 that $E(\sigma)$ satisfies some timed property. A related problem, shield synthesis [5], asks given a
 96 network of deterministic I/O timed automata \mathcal{N} that communicate with their environment, to
 97 synthesize two additional components, a pre-shield, that reads outputs from the environment
 98 and produces inputs for \mathcal{N} , and a post-shield, that reads outputs from \mathcal{N} and produces
 99 outputs to the environment to satisfy timed safety properties when faults (timing, location
 100 errors,...) occur. Synthesis is achieved using timed games. Unlike these, our goal is not to
 101 avoid violation of a property, but rather to verify that the system *recovers within boundedly*
 102 *many steps*, from a possibly large time deviation w.r.t. its behavior. Finally, *faults* in timed
 103 automata have also been studied in a diagnosis setting, e.g. in [6], where faults are detected
 104 within a certain delay from partial observation of runs.

105 2 Preliminaries

106 Let Σ be a finite non-empty alphabet and $\Sigma^\infty = \Sigma^* \cup \Sigma^\omega$ a set of finite or infinite words over
 107 Σ . $\mathbb{R}, \mathbb{R}_{\geq 0}, \mathbb{Q}, \mathbb{N}$ respectively denote the set of real numbers, non-negative reals, rationals,
 108 and natural numbers. We write $(\Sigma \times \mathbb{R}_{\geq 0})^\infty = (\Sigma \times \mathbb{R}_{\geq 0})^* \cup (\Sigma \times \mathbb{R}_{\geq 0})^\omega$ for finite or infinite
 109 timed words over Σ . A finite (infinite) timed word has the form $w = (a_1, d_1) \dots (a_n, d_n)$
 110 (resp. $w = (a_1, d_1) \dots$) where for every i , $d_i \leq d_{i+1}$. For $i \leq j$, we denote by $w_{[i,j]}$, the
 111 sequence $(a_i, d_i) \dots (a_j, d_j)$. *Untiming* of a timed word $w \in (\Sigma \times \mathbb{R}_{\geq 0})^\infty$ denoted $Unt(w)$, is
 112 its projection on the first component, and is a word in Σ^∞ . A *clock* is a real-valued variable x
 113 and an *atomic clock constraint* is an inequality of the form $a \bowtie_l x \bowtie_u b$, with $\bowtie_l, \bowtie_u \in \{\leq, <\}$,
 114 $a, b \in \mathbb{N}$. An atomic *diagonal constraint* is of the form $a \bowtie_l x - y \bowtie_u b$, where x and y are
 115 different clocks. *Guards* are conjunctions of atomic constraints on a set X of clocks.

116 ► **Definition 1.** A *timed automaton*[2] is a tuple $\mathcal{A} = (L, I, X, \Sigma, T, F)$ with finite set of
 117 locations L , initial locations $I \subseteq L$, finitely many clocks X , finite action set Σ , final locations
 118 $F \subseteq L$, and transition relation $T \subseteq L \times \mathcal{G} \times \Sigma \times 2^X \times L$ where \mathcal{G} are guards on X .

119 A *valuation* of a set of clocks X is a map $\nu : X \rightarrow \mathbb{R}_{\geq 0}$ that associates a non-negative real
 120 value to each clock in X . For every clock x , $\nu(x)$ has an integral part $\lfloor \nu(x) \rfloor$ and a fractional
 121 part $\text{frac}(\nu(x)) = \nu(x) - \lfloor \nu(x) \rfloor$. We will say that a valuation ν on a set of clocks X satisfies
 122 a guard g , denoted $\nu \models g$ if and only if replacing every $x \in X$ by $\nu(x)$ in g yields a tautology.
 123 We will denote by $[g]$ the set of valuations that satisfy g . Given $\delta \in \mathbb{R}_{\geq 0}$, we denote by $\nu + \delta$
 124 the valuation that associates value $\nu(x) + \delta$ to every clock $x \in X$. A *configuration* is a pair
 125 $C = (l, \nu)$ of a location of the automaton and valuation of its clocks. The semantics of a
 126 timed automaton is defined in terms of discrete and timed moves from a configuration to the
 127 next one. A *timed move of duration* δ lets $\delta \in \mathbb{R}_{\geq 0}$ time units elapse from a configuration
 128 $C = (l, \nu)$ which leads to configuration $C' = (l, \nu + \delta)$. A *discrete move* from configuration
 129 $C = (l, \nu)$ consists of taking one of the transitions leaving l , i.e., a transition of the form
 130 $t = (l, g, a, R, l')$ where g is a guard, $a \in \Sigma$ a particular action name, R is the set of clocks
 131 reset by the transition, and l' the next location reached. A discrete move with transition t is
 132 allowed only if $\nu \models g$. Taking transition t leads the automaton to configuration $C' = (l', \nu')$
 133 where $\nu'(x) = \nu(x)$ if $x \notin R$, and $\nu'(x) = 0$ otherwise.

134 ► **Definition 2** (Runs, Maximal runs, Accepting runs). *An (infinite) run of a timed automaton*
 135 \mathcal{A} *is a sequence* $\rho = (l_0, \nu_0) \xrightarrow{(t_1, d_1)} (l_1, \nu_1) \xrightarrow{(t_2, d_2)} \dots$ *where every pair* (l_i, ν_i) *is a configuration,*
 136 *and there exists an (infinite) sequence of timed and discrete moves* $\delta_1.t_1.\delta_2.t_2\dots$ *in* \mathcal{A} *such*
 137 *that* $\delta_i = d_{i+1} - d_i$, *and a timed move of duration* δ_i *from* (l_i, ν_i) *to* $(l_i, \nu_i + \delta_i)$ *and a discrete*
 138 *move from* $(l_i, \nu_i + \delta_i)$ *to* (l_{i+1}, ν_{i+1}) *via transition* t_i . *A run is maximal if it is infinite, or if*
 139 *it ends at a location with no outgoing transitions. A finite run is accepting if its last location*
 140 *is final, while an infinite run is accepting if it visits accepting locations infinitely often.*

141 One can associate a finite/infinite *timed word* w_ρ to every run ρ of \mathcal{A} by letting $w_\rho = (a_1, d_1)$
 142 $(a_2, d_2) \dots (a_n, d_n) \dots$, where a_i is the action in transition t_i and d_i is the time stamp of
 143 t_i in ρ . A (finite/infinite) *timed word* w is accepted by \mathcal{A} if there exists a (finite/infinite)
 144 accepting run ρ such that $w = w_\rho$. The *timed language* of \mathcal{A} is the set of all timed words
 145 accepted by \mathcal{A} , and is denoted by $\mathcal{L}(\mathcal{A})$. The *untimed language* of \mathcal{A} is the language
 146 $Unt(\mathcal{L}(\mathcal{A})) = \{Unt(w) \mid w \in \mathcal{L}(\mathcal{A})\}$. As shown in [2], the untimed language of a timed
 147 automaton can be captured by an abstraction called the *region automaton*. Formally, given
 148 a clock x , let c_x be the largest constant in an atomic constraint of a guard of \mathcal{A} involving x .
 149 Two valuations ν, ν' of clocks in X are *equivalent*, written $\nu \sim \nu'$ if and only if:

- 150 i) $\forall x \in X$, either $\lfloor \nu(x) \rfloor = \lfloor \nu'(x) \rfloor$ or both $\nu(x) \geq c_x$ and $\nu'(x) \geq c_x$
- 151 ii) $\forall x, y \in X$ with $\nu(x) \leq c_x$ and $\nu(y) \leq c_y$, $\text{frac}(\nu(x)) \leq \text{frac}(\nu(y))$ iff $\text{frac}(\nu'(x)) \leq \text{frac}(\nu'(y))$
- 152 iii) For all $x \in X$ with $\nu(x) \leq c_x$, $\text{frac}(\nu(x)) = 0$ iff $\text{frac}(\nu'(x)) = 0$.

153 A *region* r of \mathcal{A} is the equivalence class induced by \sim . For a valuation ν , we denote by $[\nu]$
 154 the region of ν , i.e., its equivalence class. We will also write $\nu \in r$ (ν is a valuation in region r
 155 when $r = [\nu]$). For a given automaton \mathcal{A} , there exists only a finite number of regions, bounded
 156 by 2^K , where K is the size of the constraints set in \mathcal{A} . It is well known that for a clock
 157 constraint ψ that, if $\nu \sim \nu'$, then $\nu \models \psi$ if and only if $\nu' \models \psi$. A region r' is a *time successor*
 158 of another region r if for every $\nu \in r$, there exists $\delta \in \mathbb{R}_{>0}$ such that $\nu + \delta \in r'$. We denote by
 159 $Reg(X)$ the set of all possible regions of the set of clocks X . A region r satisfies a guard g if
 160 and only if there exists a valuation $\nu \in r$ such that $\nu \models g$. The region automaton of a timed
 161 automaton $\mathcal{A} = (L, I, X, \Sigma, T, F)$ is the untimed automaton $\mathcal{R}(\mathcal{A}) = (S_R, I_R, \Sigma, T_R, F_R)$ that
 162 recognizes the untimed language $Unt(\mathcal{L}(\mathcal{A}))$. States of $\mathcal{R}(\mathcal{A})$ are of the form (l, r) , where l is a
 163 location of \mathcal{A} and r a region, i.e., $S_R \subseteq L \times Reg(X)$, $I_R \subseteq I \times Reg(X)$, and $F_R \subseteq F \times Reg(X)$.
 164 The transition relation T_R is such that $((l, r), a, (l', r')) \in T_R$ if there exists a transition
 165 $t = (l, g, a, R, l') \in T$ such that there exists a time successor region r'' of r such that r''
 166 satisfies the guard g , and r' is obtained from r'' by resetting values of clocks in R . The size of
 167 the region automaton is the number of states in $\mathcal{R}(\mathcal{A})$ and is denoted $|\mathcal{R}(\mathcal{A})|$. For a region r
 168 defined on a set of clocks Y , we define a projection operator $\Pi_X(r)$ to represent the region r
 169 projected on the set of clocks $X \subseteq Y$. Let $\rho = (l_0, \nu_0) \xrightarrow{(t_1, d_1)} (l_1, \nu_1) \dots$ be a run of \mathcal{A} , where
 170 every t_i is of the form $t_i = (l_i, g_i, a_i, R_i, l'_i)$. The *abstract run* $\sigma_\rho = (l_0, r_0) \xrightarrow{a_1} (l_1, r_1) \dots$ of ρ
 171 is a path in the region automata $\mathcal{R}(\mathcal{A})$ such that, $\forall i \in \mathbb{N}, r_i = [\nu_i]$. We represent runs using
 172 variables ρ, π and the corresponding abstract runs with σ_ρ, σ_π respectively. The region
 173 automaton can be used to prove non-emptiness, as $\mathcal{L}(\mathcal{A}) \neq \emptyset$ iff $\mathcal{R}(\mathcal{A})$ accepts some word.

3 Resilience Problems

174 We define the semantics of timed automata when perturbations can delay the occurrence
 175 of an action. Consider a transition $t = (l, g, a, R, l')$, with $g ::= x \leq 10$, where action a can
 176 occur as long as x has not exceeded 10. Timed automata have a idealized representation of
 177 time, and do not consider perturbations that occur in real systems. Consider, for instance
 178 that ‘ a ’ is a physical event planned to occur at a maximal time stamp 10: a water tank
 179

reaches its maximal level, a train arrives in a station etc. These events can be delayed, and nevertheless occur. One can even consider that uncontrollable delays are part of the normal behavior of the system, and that $\mathcal{L}(\mathcal{A})$ is the ideal behavior of the system, when all delays are met. In the rest of the paper, we propose a fault model that assigns a maximal error to each fireable action. This error model is used to encode the fact that an action might occur at time points slightly greater than what is allowed in the original model semantics.

► **Definition 3 (Fault model).** A fault model \mathcal{P} is a map $\mathcal{P} : \Sigma \rightarrow \mathbb{Q}_{\geq 0}$ that associates to every action in $a \in \Sigma$ a possible maximal delay $\mathcal{P}(a) \in \mathbb{Q}_{\geq 0}$.

For simplicity, we consider only executions in which a single timing error occurs. The perturbed semantics defined below easily adapts to a setting with multiple timing errors. With a fault model, we can define a new timed automaton, for which every run $\rho = (l_0, \nu_0) \xrightarrow{(t_1, d_1)} (l_1, \nu_1) \xrightarrow{(t_2, d_2)} \dots$ contains at most one transition $t_i = (l, g, a, r, l')$ occurring later than allowed by guard g , and agrees with a run of \mathcal{A} until this faulty transition is taken.

► **Definition 4 (Enlargement of a guard).** Let ϕ be an inequality of the form $a \bowtie_l x \bowtie_u b$, where $\bowtie_l, \bowtie_u \in \{\leq, <\}$. The enlargement of ϕ by a time error δ is the inequality $\phi_{\triangleright\delta}$ of the form $a \bowtie_l x \leq b + \delta$. Let g be a guard of the form

$$g = \bigwedge_{i \in 1..m} \phi_i = a_i \bowtie_{l_i} x_i \bowtie_{u_i} b_i \wedge \bigwedge_{j \in 1..q} \phi_j = a_j \bowtie_{l_j} x_j - y_j \bowtie_{u_j} b_j.$$

The enlargement of g by δ is the guard $g_{\triangleright\delta} = \bigwedge_{i \in 1..m} \phi_{i \triangleright\delta} \wedge \bigwedge_{j \in 1..q} \phi_j$

For every transition $t = (l, g, a, R, l')$ with enlarged guard

$$g_{\triangleright\mathcal{P}(a)} = \bigwedge_{i \in 1..m} \phi_i = a_i \bowtie_{l_i} x_i \leq b_i + \mathcal{P}(a) \wedge \bigwedge_{j \in 1..q} \phi_j = a_j \bowtie_{l_j} x_j - y_j \bowtie_{u_j} b_j,$$

we can create a new transition $t_{f, \mathcal{P}} = (l, g_{f, \mathcal{P}}, a, R, l')$ called a faulty transition such that,

$$g_{f, \mathcal{P}} = \bigwedge_{i \in 1..m} \phi_i = b_i \bowtie_{l_i} x_i \leq b_i + \mathcal{P}(a) \wedge \bigwedge_{j \in 1..q} \phi_j = a_j \bowtie_{l_j} x_j - y_j \bowtie_{u_j} b_j \text{ with } \bowtie_{l_i} \in \{<, \leq\} \setminus \bowtie_{l_i}$$

Diagonal constraints remain unchanged under enlargement, as the difference between clocks x and y is preserved on time elapse. From now, we fix a fault model \mathcal{P} and write t_f and g_f instead of $t_{f, \mathcal{P}}$ and $g_{f, \mathcal{P}}$. Clearly, g and g_f are disjoint, and $g \vee g_f$ is equivalent to $g_{\triangleright\delta}$.

► **Definition 5 (Enlargement of automata).** Let $\mathcal{A} = (L, I, X, \Sigma, T, F)$ be a timed automaton. The enlargement of \mathcal{A} by a fault model \mathcal{P} is the automaton $\mathcal{A}_{\mathcal{P}} = (L_{\mathcal{P}}, I, X, \Sigma, T_{\mathcal{P}}, F_{\mathcal{P}})$, where

■ $L_{\mathcal{P}} = L \cup \{\dot{l} \mid l \in L\}$ and $F_{\mathcal{P}} = F \cup \{\dot{l} \mid l \in F\}$. A location \dot{l} indicates that an unexpected delay has occurred.

■ $T_{\mathcal{P}} = T \cup \dot{T}$ such that, $\dot{T} = \{(l, g_f, a, R, \dot{l}') \mid (l, g, a, R, l') \in T\} \cup \{(\dot{l}, g, a, R, \dot{l}') \mid (l, g, a, R, l') \in T\}$ i.e., \dot{T} is the set of transitions occurring after a fault.

A run of $\mathcal{A}_{\mathcal{P}}$ is faulty if it contains a transition of \dot{T} . It is just faulty if its last transition belongs to \dot{T} and all other transitions belong to T . Note that while faulty runs can be finite or infinite, just faulty runs are always finite prefix of a faulty run, and end in a location \dot{l} .

► **Definition 6 (Back To Normal (BTN)).** Let $K \geq 1$, \mathcal{A} be a timed automaton with fault model \mathcal{P} . Let $\rho = (l_0, \nu_0) \xrightarrow{(t_1, d_1)} (l_1, \nu_1) \xrightarrow{(t_2, d_2)} \dots$ be a (finite or infinite) faulty accepting run of $\mathcal{A}_{\mathcal{P}}$, with associated timed word $(a_1, d_1)(a_2, d_2) \dots$ and let $i \in \mathbb{N}$ be the position of the faulty transition in ρ . Then ρ is back to normal (BTN) after K steps if there exists an accepting run $\rho' = (l'_0, \nu'_0) \xrightarrow{(t'_1, d'_1)} (l'_1, \nu'_1) \xrightarrow{(t'_2, d'_2)} \dots$ of \mathcal{A} with associated timed word $(a'_1, d'_1)(a'_2, d'_2) \dots$ and an index $\ell \in \mathbb{N}$ such that $(a'_\ell, d'_\ell)(a'_{\ell+1}, d'_{\ell+1}) \dots = (a_{i+K}, d_{i+K})(a_{i+K+1}, d_{i+K+1}) \dots \rho$

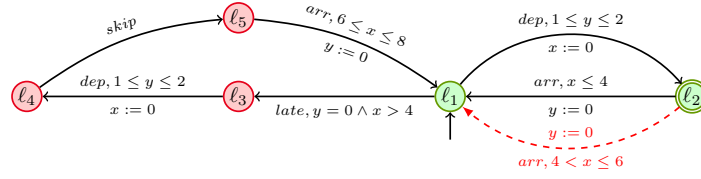


Figure 1 Model of a train system with a mechanism to recover from delays

is untimed back to normal (untimed BTN) after K steps if there exists an accepting run $\rho' = (l'_0, \nu'_0) \xrightarrow{(t'_1, d'_1)} (l'_1, \nu'_1) \xrightarrow{(t'_2, d'_2)} \dots$ of \mathcal{A} and an index $\ell \in \mathbb{N}$ s.t. $a'_\ell a'_{\ell+1} \dots = a_{i+K} a_{i+K+1} \dots$

In other words, if w is a timed word having a faulty accepting run (i.e., $w \in \mathcal{L}(\mathcal{A}_P)$), the suffix of w , K steps after the fault, matches with the suffix of some word $w' \in \mathcal{L}(\mathcal{A})$. Note that the accepting run of w' in \mathcal{A} is not faulty, by definition. The conditions in untimed BTN are simpler, and ask the same sequence of actions, but not equality on dates.

Our current definition of back-to-normal in K steps means that a system recovered from a fault (a primary delay) in $\leq K$ steps and remained error-free. We can generalize our definition, to model real life situations where more than one fault happens due to time delays, but the system recovers from each one in a small number of steps and eventually achieves its fixed goal (a reachability objective, some ω -regular property...). A classical example of this is a metro network, where trains are often delayed, but nevertheless recover from these delays to reach their destination on time. This motivates the following definition of resilience.

- **Definition 7 (Resilience).** A timed automaton \mathcal{A} is
 - (untimed) K - \forall -resilient if every finite faulty accepting run is (untimed) BTN in K steps.
 - (untimed) K - \exists -resilient if every just faulty run ρ_{jf} can be extended into a maximal accepting run ρ_f which is (untimed) BTN in K steps.

Intuitively, a faulty run of \mathcal{A} is BTN if the system has definitively recovered from a fault, i.e., it has recovered and will follow the behavior of the original system after its recovery. The definition of existential resilience considers maximal (infinite, or finite but ending at a location with no outgoing transitions) runs to avoid situations where an accepting faulty run ρ_f is BTN, but all its extensions i.e., suffixes $\rho_f \cdot \rho'$ are such that $\rho_f \cdot \rho'$ is not BTN.

► **Example 8.** We model train services to a specific destination such as an airport. On an average, the distance between two consecutive stations is covered in ≤ 4 time units. At each stop in a station, the dwell time is in between 1 and 2 time units. To recover from a delay, the train is allowed to skip an intermediate station (as long as the next stop is not the destination). Skipping a station is a choice, and can only be activated if there is a delay. We model this system with the timed automaton of Figure 1. There are 5 locations: ℓ_1 , and ℓ_2 represent the normal behavior of the train and ℓ_3, ℓ_4, ℓ_5 represent the skipping mechanism. These locations can only be accessed if the faulty transition (represented as a red dotted arrow in Figure 1) is fired. A transition t_{ij} goes from ℓ_i to ℓ_j , and t_{21} denotes the faulty transition from ℓ_2 to ℓ_1 . The green locations represent the behavior of the train without any delay, and the red locations represent behaviors when the train chooses to skip the next station to recover from a delay. This mechanism is invoked once the train leaves the station where it arrived late (location ℓ_3). When it departs, x is reset as usual; the next arrival to a station (from location ℓ_4) happens after skipping stop at the next station. The delay can be recovered since the running time since the last stop (covering 2 stations) is between 6 and 8 units of time. Asking if this system is resilient amounts to asking if this mechanism can be used to recover from delays (defined by the fault model $\mathcal{P}(arr) = 2$,

see Figure 3, Appendix A). Consider the faulty run $\rho_f = (\ell_1, 0|0) \xrightarrow{(t_{12},2)} (\ell_2, 0|2) \xrightarrow{(t_{21},8)}$
 $(\ell_1, 6|0) \xrightarrow{(t_{13},8)} (\ell_3, 6|0) \xrightarrow{(t_{34},10)} (\ell_4, 0|2) \xrightarrow{(t_{45},10)} (\ell_5, 0|2) \xrightarrow{(t_{51},18)} (\ell_1, 8|0) \xrightarrow{(t_{12},19)} (\ell_2, 0|1)$ reading
 $(dep, 2)(arr, 8)(late, 8)(dep, 10)(skip, 10)(arr, 18)(dep, 19)$. Run ρ_f is BTN in 4 steps. It
 matches the non-faulty run $\rho = (\ell_1, 0|0) \xrightarrow{(t_{12},2)} (\ell_2, 0|2) \xrightarrow{(t_{21},6)} (\ell_1, 4|0) \xrightarrow{(t_{12},8)} (\ell_2, 0|2) \xrightarrow{(t_{21},12)}$
 $(\ell_1, 4|0) \xrightarrow{(t_{12},14)} (\ell_2, 0|2) \xrightarrow{(t_{21},18)} (\ell_1, 4|0) \xrightarrow{(t_{12},19)} (\ell_2, 0|1)$ reading $(dep, 2)(arr, 6)(dep, 8)(arr, 12)$
 $(dep, 14)(arr, 18)(dep, 19)$. This automaton is K - \exists -resilient for $K = 4$ and fault model \mathcal{P} ,
 as skipping a station after a delay of ≤ 2 time units allows to recover the time lost. It is
 not K - \forall -resilient, for any K , as skipping is not mandatory, and a train can be late for an
 arbitrary number of steps. In Appendix A we give another example that is 1- \forall -resilient.

K - \forall -resilience always implies K - \exists -resilience. In case of K - \forall -resilience, every faulty run
 ρ_w has to be BTN in $\leq K$ steps after the occurrence of a fault. This implies K - \exists -resilience
 since, any just faulty run ρ_w that is the prefix of an accepting run ρ of $\mathcal{A}_{\mathcal{P}}$ is BTN in less
 than K steps. The converse does not hold: $\mathcal{A}_{\mathcal{P}}$ can have a pair of runs ρ_1, ρ_2 , sharing a
 common just faulty run ρ_f as prefix such that ρ_1 is BTN in K steps, witnessing existential
 resilience, while ρ_2 is not. Finally, an accepting run $\rho = \rho_f \rho_s$ in $\mathcal{A}_{\mathcal{P}}$ s.t., ρ_f is just faulty
 and $|\rho_s| < K$, is BTN in K steps since ε is a suffix of a run accepted by \mathcal{A} .

4 Existential Resilience

In this section, we consider existential resilience both in the timed and untimed settings.

Existential Timed Resilience. As the first step, we define a product automaton $\mathcal{B} \otimes_K \mathcal{A}$
 that recognizes BTN runs. Intuitively, the product synchronizes runs of \mathcal{B} and \mathcal{A} as soon as
 \mathcal{B} has performed K steps after a fault, and guarantees that actions performed by \mathcal{A} and \mathcal{B}
 are performed at the same date in the respective runs of \mathcal{A} and \mathcal{B} . Before this synchronization,
 \mathcal{A} and \mathcal{B} take transitions or stay in the same location, but let the same amount of time
 elapse, guaranteeing that synchronization occurs after runs of \mathcal{A} and \mathcal{B} of identical durations.
 The only way to ensure this with a timed automaton is to track the global timing from the
 initial state of both automata \mathcal{A} and \mathcal{B} till K steps after the fault, even though we do not
 need the timing for individual actions till K steps after the fault.

► **Definition 9 (Product).** Let $\mathcal{A} = (L_A, I_A, X_A, \Sigma, T_A, F_A)$ and $\mathcal{B} = (L_B, I_B, X_B, \Sigma, T_B, F_B)$
 be two timed automata, where \mathcal{B} contains faulty transitions. Let $K \in \mathbb{N}$ be an integer. Then,
 the product $\mathcal{B} \otimes_K \mathcal{A}$ is a tuple $(L, I, X_A \cup X_B, (\Sigma \cup \{*\})^2, T, F)$ where $L \subseteq \{L_B \times L_A \times [-1, K]\}$,
 $F = L_B \times F_A \times [-1, K]$, and initial set of states $I = I_B \times I_A \times \{-1\}$. Intuitively, -1 means
 no fault has occurred yet. Then we assign K and decrement to 0 to denote that K steps after
 fault have passed. The set of transitions T is as follows: We have $((l_B, l_A, n), g, \langle x, y \rangle$
 $, R, (l'_B, l'_A, n')) \in T$ if and only if either:

- $n \neq 0$ (no fault has occurred, or less than K steps of \mathcal{B} have occurred), the action is
 $\langle x, y \rangle = \langle a, * \rangle$, we have transition $t_B = (l_B, g, a, R, l'_B) \in T_B$, $l_A = l'_A$ (the location
 of \mathcal{A} is unchanged) and either: $n = -1$, the transition t_B is faulty and $n' = K$, or $n = -1$,
 the transition t_B is non faulty and $n' = -1$, or $n > 0$ and $n' = n - 1$.
- $n = n' \neq 0$ (no fault has occurred, or less than K steps of \mathcal{B} have occurred), the action
 is $\langle x, y \rangle = \langle *, a \rangle$, we have the transition $t_A = (l_A, g, a, R, l'_A) \in T_A$, $l_B = l'_B$ (the
 location of \mathcal{B} is unchanged).
- $n = n' = 0$ (at least K steps after a fault have occurred), the action is $\langle x, y \rangle = \langle a, a \rangle$
 and there exists two transitions $t_B = (l_B, g, a, R_B, l'_B) \in T_B$ and $t_A = (l_A, g_A, a, R_A, l'_A) \in$
 T_A with $g = g_A \wedge g_B$, and $R = R_B \cup R_A$ (t_A and t_B occur synchronously).

303 Runs of $\mathcal{B} \otimes_K \mathcal{A}$ are sequences of the form $\rho^\otimes = (l_0, l_0^A, n_0) \xrightarrow{(t_1, t_1^A), d_1} \dots \xrightarrow{(t_k, t_k^A), d_k} (l_k, l_k^A, n_k)$
 304 where each $(t_i, t_i^A) \in (T_B \cup \{t_*\}) \times (T_A \cup \{t_*^A\})$ defines uniquely the transition of $\mathcal{B} \otimes_K \mathcal{A}$,
 305 where t_* corresponds to the transitions with action $*$. Transitions are of types (t_i, t_i^A) or
 306 (t_*, t_*^A) up to a fault and K steps of T_B , and $(t_i, t_i^A) \in T_B \times T_A$ from there on.

307 For any timed run ρ^\otimes of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$, the projection of ρ^\otimes on its first component is a timed
 308 run ρ of $\mathcal{A}_{\mathcal{P}}$, that is projecting ρ^\otimes on transitions of $\mathcal{A}_{\mathcal{P}}$ and remembering only location and
 309 clocks of $\mathcal{A}_{\mathcal{P}}$ in states. In the same way, the projection of ρ^\otimes on its second component is a
 310 timed run ρ' of \mathcal{A} . Given timed runs ρ of $\mathcal{A}_{\mathcal{P}}$ and ρ' of \mathcal{A} , we denote by $\rho \otimes \rho'$ the timed
 311 run (if it exists) of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ such that the projection on the first component is ρ and the
 312 projection on the second component is ρ' . For $\rho \otimes \rho'$ to exist, we need ρ, ρ' to have the same
 313 duration, and for ρ_s the suffix of ρ starting K steps after a fault (if there is a fault and K
 314 steps, $\rho_s = \varepsilon$ the empty run otherwise), ρ_s needs to be suffix of ρ' as well.

315 A run ρ^\otimes of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ is accepting if its projection on the second component (\mathcal{A}) is
 316 accepting (i.e., ends in an accepting state if it is finite and goes through an infinite number
 317 of accepting state if it is infinite). We can now relate the product $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ to BTN runs.

318 ► **Proposition 10.** *Let ρ_f be a faulty accepting run of $\mathcal{A}_{\mathcal{P}}$. The following are equivalent:*

- 319 i ρ_f is BTN in K -steps
- 320 ii there is an accepting run ρ^\otimes of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ s.t., the projection on its first component is ρ_f

321 Let ρ be a finite run of $\mathcal{A}_{\mathcal{P}}$. We denote by $T_\rho^{\otimes K}$ the set of states of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ such that
 322 there exists a run ρ^\otimes of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ ending in this state, whose projection on the first component
 323 is ρ . We then define $S_\rho^{\otimes K}$ as the set of states of $\mathcal{R}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A})$ corresponding to $T_\rho^{\otimes K}$, i.e.,
 324 $S_\rho^{\otimes K} = \{(s, [\nu]) \in \mathcal{R}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}) \mid (s, \nu) \in T_\rho^{\otimes K}\}$. If we can compute the set $\mathbb{S} = \{S_\rho^{\otimes K} \mid \rho$
 325 is a finite run of $\mathcal{A}_{\mathcal{P}}\}$, we would be able to solve *timed* universal resilience. Proposition 18
 326 shows that universal resilience is undecidable. Hence, computing \mathbb{S} is impossible. Roughly
 327 speaking, it is because this set depends on the exact timing in a run ρ , and in general one
 328 cannot use the region construction.

329 We can however show that in some restricted cases, we can use a *modified* region
 330 construction to build $S_\rho^{\otimes K}$, which will enable decidability of timed existential resilience. First,
 331 we restrict to *just faulty runs*, i.e., consider runs of $\mathcal{A}_{\mathcal{P}}$ and \mathcal{A} of equal durations, but that
 332 did not yet synchronize on actions in the product $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$. For a timed run ρ , by its
 333 duration, we mean the time-stamp or date of occurrence of its last event. Second, we consider
 334 abstract runs $\tilde{\sigma}$ through a so-called *strong region automaton*, as defined below. Intuitively, $\tilde{\sigma}$
 335 keeps more information than in the usual region automaton to ensure that for two timed
 336 runs $\rho_1 = (t_1, d_1)(t_2, d_2) \dots$, and $\rho_2 = (t_1, e_1)(t_2, e_2) \dots$ associated with the same run of
 337 the strong region automaton, we have $[e_i] = [d_i]$ for all i . Formally, we build the strong
 338 region automaton $\mathcal{R}_{\text{strong}}(\mathcal{B})$ of a timed automaton \mathcal{B} as follows. We add a virtual clock
 339 x to \mathcal{B} which is reset at each integral time point, add constraint $x < 1$ to each transition
 340 guard, and add a virtual self loop transition with guard $x = 1$ resetting x on each state.
 341 We then make the usual region construction on this extended timed automaton to obtain
 342 $\mathcal{R}_{\text{strong}}(\mathcal{B})$. The strong region construction thus has the same complexity as the standard
 343 region construction. Let $\mathcal{L}(\mathcal{R}_{\text{strong}}(\mathcal{B}))$ be the language of this strong region automaton, where
 344 these self loops on the virtual clock are projected away. Indeed, these additional transitions,
 345 added to capture ticks at integral times, do not change the overall behavior of \mathcal{B} , i.e., we
 346 have $Unt(\mathcal{L}(\mathcal{B})) \subseteq \mathcal{L}(\mathcal{R}_{\text{strong}}(\mathcal{B})) \subseteq \mathcal{L}(\mathcal{R}(\mathcal{B})) = Unt(\mathcal{L}(\mathcal{B}))$ so $Unt(\mathcal{L}(\mathcal{B})) = \mathcal{L}(\mathcal{R}_{\text{strong}}(\mathcal{B}))$.

347 For a finite abstract run $\tilde{\sigma}$ of the *strong* region automaton $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})$, we define the set
 348 $S_{\tilde{\sigma}}^{\otimes K}$ of states of $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A})$ (the virtual clock is projected away, and our region is
 349 wrt original clocks) such that there exists a run $\tilde{\sigma}^\otimes$ through $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A})$ ending in this

350 state and whose projection on the first component is $\tilde{\sigma}$. Let $\tilde{\sigma}_\rho$ be the run of $\mathcal{R}_{\text{strong}}(\mathcal{A}_P)$
 351 associated with a run ρ of \mathcal{A}_P . It is easy to see that $S_{\tilde{\sigma}}^{\otimes K} = \bigcup_{\rho|\tilde{\sigma}_\rho=\tilde{\sigma}} S_{\rho}^{\otimes K}$. For a *just faulty*
 352 timed run ρ of \mathcal{A}_P , we have a stronger relation between $S_{\rho}^{\otimes K}$ and $S_{\tilde{\sigma}_\rho}^{\otimes K}$:

353 ► **Proposition 11.** *Let ρ be a just faulty run of \mathcal{A}_P . Then $S_{\rho}^{\otimes K} = S_{\tilde{\sigma}_\rho}^{\otimes K}$.*

354 **Proof.** First, notice that given a just faulty timed run ρ of \mathcal{A}_P and a timed run ρ' of \mathcal{A}
 355 same duration, the timed run $\rho \otimes \rho'$ (the run of $\mathcal{A}_P \otimes_K \mathcal{A}$ such that ρ is the projection on
 356 the first component and ρ' on the second component) exists.

357 To show that $S_{\rho}^{\otimes K} = S_{\tilde{\sigma}_\rho}^{\otimes K}$, we show that for any pair of just faulty runs ρ_1, ρ_2 of \mathcal{A}_P with
 358 $\tilde{\sigma}_{\rho_1} = \tilde{\sigma}_{\rho_2}$, we have $S_{\rho_1}^{\otimes K} = S_{\rho_2}^{\otimes K}$, which yields the result as $S_{\tilde{\sigma}_\rho}^{\otimes K} = \bigcup_{\rho'|\tilde{\sigma}_{\rho'}=\tilde{\sigma}_\rho} S_{\rho'}^{\otimes K}$. Consider
 359 ρ_1, ρ_2 , two just faulty timed runs of \mathcal{A}_P with $\tilde{\sigma}_{\rho_1} = \tilde{\sigma}_{\rho_2}$ and let $(l_{\mathcal{A}_P}, l_A, K, r) \in S_{\rho_1}^{\otimes K}$. Then,
 360 this implies that there exists $\nu_1 \models r$ and a timed run ρ'_1 of \mathcal{A} with the same duration as ρ_1 ,
 361 such that $\rho_1 \otimes \rho'_1$ ends in state $(l_{\mathcal{A}_P}, l_A, K, \nu_1)$. The following lemma completes the proof:

362 ► **Lemma 12.** *There exists $\nu_2 \models r$ and a timed run ρ'_2 of \mathcal{A} with the same duration as ρ_2 ,
 363 such that $\rho_2 \otimes \rho'_2$ ends in state $(l_{\mathcal{A}_P}, l_A, K, \nu_2)$.*

364 The idea of the proof (detailed in appendix B) is to show that we can construct ρ'_2
 365 which will have the same transitions as ρ'_1 , with same integral parts in timings (thanks to
 366 the information from the strong region automaton), but possibly different timings in the
 367 fractional parts, called a retiming of ρ'_1 . Notice that ρ_2 is a retiming of ρ_1 , as $\tilde{\sigma}_{\rho_1} = \tilde{\sigma}_{\rho_2}$. We
 368 translate the requirement on ρ'_2 into a set of constraints (which is actually a partial ordering)
 369 on the fractional parts of the dates of its transitions, and show that we can indeed set the
 370 dates accordingly. This translation follows the following idea: the value of a clock x just
 371 before firing transition t is obtained by considering the date d of t minus the date d^x of the
 372 latest transition t^x at which x has been last reset before t . In particular, the difference $x - y$
 373 between clocks x, y just before firing transition t is $(d - d^x) - (d - d^y) = d^y - d^x$. That is, the
 374 value of a clock or its difference can be obtained by considering the difference between two
 375 dates of transitions. A constraint given by $x - y \in (n, n + 1)$ is equivalent with the constraint
 376 given by $d^y - d^x \in (n, n + 1)$, and similar constraints on the fractional parts can be given.

377 Lemma 12 gives our result immediately. Indeed, the lemma implies that $(l_{\mathcal{A}_P}, l_A, K, r) \in$
 378 $S_{\rho_2}^{\otimes K}$ from which we infer that $S_{\rho_1}^{\otimes K} \subseteq S_{\rho_2}^{\otimes K}$. By a symmetric argument we get the other
 379 containment also, and hence we conclude that $S_{\rho_1}^{\otimes K} = S_{\rho_2}^{\otimes K}$. ◀

380 **Algorithm to solve Existential Timed Resilience.** We can now consider existential
 381 timed resilience, and prove that it is decidable thanks to Propositions 10 and 11. The
 382 main idea is to reduce the existential resilience question to a question on the sets of regions
 383 reachable after just faulty runs. Indeed, focusing on just faulty runs means that we do not
 384 have any actions to match, only the duration of the run till the fault, whereas if we had tried
 385 to reason on faulty runs in general, actions have to be synchronized K steps after the fault
 386 and then we cannot compute the set of $S_{\rho_f}^{\otimes K}$. We can show that reasoning on $S_{\rho_f}^{\otimes K}$ for just
 387 faulty runs is sufficient. Let ρ_f be a just faulty timed run of \mathcal{A}_P . We say that $s \in S_{\rho_f}^{\otimes K}$ is
 388 *safe* if there exists a (finite or infinite) maximal accepting run of $\mathcal{A}_P \otimes_K \mathcal{A}$ from s , and that
 389 $S_{\rho_f}^{\otimes K}$ is safe if there exists $s \in S_{\rho_f}^{\otimes K}$ which is safe.

390 ► **Lemma 13.** *There exists a maximal accepting extension of a just faulty run ρ_f that is*
 391 *BTN in K -steps iff $S_{\rho_f}^{\otimes K}$ is safe. Further, deciding if $S_{\rho_f}^{\otimes K}$ is safe can be done in PSPACE.*

392 **Proof.** Let ρ_f a just faulty run. By Proposition 10, there exists an extension ρ of ρ_f that is
 393 BTN in K steps if and only if there exists an accepting run $\rho^{\otimes K}$ of $\mathcal{A}_P \otimes_K \mathcal{A}$ such that ρ_f

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394 is a prefix of the projection of $\rho^{\otimes K}$ on its first component, if and only if there exists a just
 395 faulty run $\rho_f^{\otimes K}$ of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ such that its projection on the first component is ρ_f , and such
 396 that an accepting state of $\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A}$ can be reached after $\rho_f^{\otimes K}$, if and only if $S_{\rho_f}^{\otimes K}$ is safe.

397 Safety of $S_{\rho_f}^{\otimes K}$ can be verified using a construction similar to the one in Theorem 16: it is
 398 hence a reachability question in a region automaton, solvable with a PSPACE complexity. ◀

399 This lemma means that it suffices to consider the set of $S_{\rho_f}^{\otimes K}$ over all ρ_f just faulty, which
 400 we can compute using region automaton thanks to Prop. 11, which gives:

401 ▶ **Theorem 14.** *K - \exists -resilience of timed automata is in EXPSPACE.*

402 **Proof.** Lemma 13 implies that \mathcal{A} is not K -timed existential resilient if and only if there exists
 403 a just faulty run ρ_f such that $S_{\rho_f}^{\otimes K}$ is not safe. This latter condition can be checked. Let us
 404 denote by $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}}) = (S_{\mathcal{R}(\mathcal{A}_{\mathcal{P}})}, I_{\mathcal{R}(\mathcal{A}_{\mathcal{P}})}, \Sigma, T_{\mathcal{R}(\mathcal{A}_{\mathcal{P}})}, F_{\mathcal{R}(\mathcal{A}_{\mathcal{P}})})$ the strong region automaton
 405 associated with $\mathcal{A}_{\mathcal{P}}$. We also denote $\mathcal{R}_{\otimes K} = (S_{\mathcal{R}_{\otimes K}}, I_{\mathcal{R}_{\otimes K}}, \Sigma, T_{\mathcal{R}_{\otimes K}}, F_{\mathcal{R}_{\otimes K}})$ the strong
 406 region automaton $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A})$. Let ρ_f be a just faulty run, and let $\sigma = \tilde{\sigma}_{\rho_f}$ denote
 407 the run of $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})$ associated with ρ_f . Thanks to Proposition 11, we have $S_{\rho_f}^{\otimes K} = S_{\sigma}^{\otimes K}$,
 408 as $S_{\rho_f}^{\otimes K}$ does not depend on the exact dates in ρ_f , but only on their regions, i.e., on σ .

409 So it suffices to find a reachable witness $S_{\sigma}^{\otimes K}$ of $\mathcal{R}_{\otimes K}$ which is not safe, to conclude that
 410 \mathcal{A} is not existentially resilient. For that, we build an (untimed) automaton \mathfrak{B} . Intuitively,
 411 this automaton follows σ up to a fault of the region automaton $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})$, and maintains
 412 the set $S_{\sigma}^{\otimes K}$ of regions of $\mathcal{R}_{\otimes K}$. This automaton stops in an accepting state immediately after
 413 occurrence of a fault. Formally, the product subset automaton \mathfrak{B} is a tuple $(S_{\mathfrak{B}}, I, \Sigma, T, F)$
 414 with set of states $S_{\mathfrak{B}} = S_{\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})} \times 2^{S_{\mathcal{R}_{\otimes K}}} \times \{0, 1\}$, set of initial states $I = I_{\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})} \times$
 415 $\{I_{\mathcal{R}_{\otimes K}}\} \times \{0\}$, and set of final states $F = S_{\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})} \times 2^{S_{\mathcal{R}_{\otimes K}}} \times \{1\}$. The set of transitions
 416 $T \subseteq S_{\mathfrak{B}} \times \Sigma \times S_{\mathfrak{B}}$ is defined as follows,

- 417 ■ $((l, r, S, 0), a, (l', r', S', b)) \in T$ if and only if $t_R = ((l, r), a, (l', r')) \in T_{\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})}$ and
 418 $b = 1$ if and only if t_R is faulty and $b = 0$ otherwise.
- 419 ■ S' is the set of states s' of $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}} \otimes_K \mathcal{A})$ whose first component is (l', r') and such
 420 that there exists $s \in S$, $(s, a, s') \in T_{\mathcal{R}(\otimes K)}$.

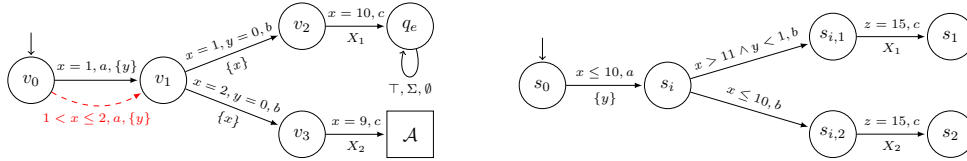
421 Intuitively, 0 in the states means no fault has occurred yet, and 1 means that a fault has
 422 just occurred, and thus no transition exists from this state. We have that for every prefix
 423 σ of a just faulty abstract run of $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})$, ending on a state (l, r) of $\mathcal{R}_{\text{strong}}(\mathcal{A}_{\mathcal{P}})$ then,
 424 there exists a unique accepting path σ^{\otimes} in \mathfrak{B} such that σ is the projection of σ^{\otimes} on its first
 425 component. Let $(l, r, S, 1)$ be the state reached by σ^{\otimes} . Then $S_{\sigma}^{\otimes K} = S$. Thus, non-existential
 426 resilience can be decided by checking reachability of a state $(l, r, S, 1)$ such that S is not
 427 safe in automaton \mathfrak{B} . As \mathfrak{B} is of doubly exponential size, reachability can be checked in
 428 EXPSPACE. Again, since EXPSPACE is closed under complementation we obtain that
 429 checking existential resilience is EXPSPACE. ◀

430 While we do not have a matching lower bound, we complete this subsection with following
 431 (easy) hardness result (we leave the details in Appendix B.1 due to lack of space).

432 ▶ **Theorem 15.** *The K - \exists -resilience problem for timed automata is PSPACE-Hard.*

433 **Existential Untimed Resilience.** We next address untimed existential resilience, which
 434 we show can be solved by enumerating states (l, r) of $\mathcal{R}(A)$ reachable after a fault, and for
 435 each of them proving existence of a BTN run starting from (l, r) . This enumeration and the
 436 following check uses polynomial space, yielding PSPACE-Completeness of K - \exists -resilience.

437 ▶ **Theorem 16.** *Untimed K - \exists -resilience is PSPACE-Complete.*



■ **Figure 2** The gadget automaton $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ (left) and the gadget \mathcal{G}_{und} (right)

438 **Proof (sketch).** *Membership* : \mathcal{A} is untimed K - \exists -resilient if and only if for all states
 439 $q = (l, r)$ reached by a just faulty run of $\mathcal{R}(\mathcal{A}_{\mathcal{P}})$, there exists a maximal accepting path σ
 440 from q such that its suffix σ_s after K steps is also the suffix of a path of $\mathcal{R}(\mathcal{A})$. This property
 441 can be verified in PSPACE. A detailed proof is provided in Appendix B.2.

442 *Hardness* : We can now show that untimed K - \exists -resilience is PSPACE-Hard. Consider a
 443 timed automaton \mathcal{A} with alphabet Σ and the construction of an automata that uses a gadget
 444 shown in Figure 2 (left). Let us call this automaton $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$. This automaton reads a word
 445 $(a, 1)(b, 1)(c, 11)$ and then accepts all timed words 2 steps after a fault, via Σ loop on a
 446 particular accepting state q_e . If $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ takes the faulty transition (marked in dotted red)
 447 then it resets all clocks of \mathcal{A} and behaves as \mathcal{A} . The accepting states are $q_e \cup F$. Then, \mathcal{A}
 448 has an accepting word if and only if $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ is untimed 2 - \exists -resilient. Since the emptiness
 449 problem for timed automata is PSPACE-Complete, the result follows. ◀

450 ▶ **Remark 17.** The hardness reduction in the proof of Theorem 16 holds even for deterministic
 451 timed automata. It is known [2] that PSPACE-Hardness of emptiness still holds for
 452 deterministic TAs. Hence, considering deterministic timed automata will not improve
 453 the complexity of K - \exists -resilience. Considering IRTAs does not change complexity either, as
 454 the gadget used in Theorem 16 can be adapted to become an IRTA (as shown in Appendix D).

455 5 Universal Resilience

456 In this section, we consider the problem of universal resilience and show that it is very close to
 457 the language inclusion question in timed automata, albeit with a few subtle differences. One
 458 needs to consider timed automata with ε -transitions [11], which are strictly more expressive
 459 than timed automata. First, we show a reduction from the language inclusion problem.

460 ▶ **Proposition 18.** *Language inclusion for timed automata can be reduced to K - \forall -resilience.*
 461 *Thus, K - \forall -resilience is undecidable in general for timed automata.*

462 **Proof Sketch.** Given $\mathcal{A}_1 = (L_1, \{l_{0_1}\}, X_1, \Sigma_1, T_1, F_1)$ and $\mathcal{A}_2 = (L_2, \{l_{0_2}\}, X_2, \Sigma_2, T_2, F_2)$,
 463 we start by defining a gadget \mathcal{G}_{und} as shown in Fig 2 (right). Next, we define an automaton
 464 \mathcal{U} that behaves as \mathcal{A}_1 after 15 time units if no fault occurs, and as \mathcal{A}_2 after 15 time units if
 465 a fault occurs. This is done by merging state s_1 in the gadget with the initial state of \mathcal{A}_1
 466 and state s_2 with the initial state of \mathcal{A}_2 . Then, we can see that $\mathcal{L}(\mathcal{A}_1) \subseteq \mathcal{L}(\mathcal{A}_2)$ if and only
 467 if \mathcal{U} is 2 - \forall -resilient. A complete proof of this theorem can be found in Appendix C. ◀

468 Next we show that the reduction is also possible in the reverse direction.

469 ▶ **Proposition 19.** *K - \forall -resilience can be reduced to language inclusion for timed automata*
 470 *with ε -transitions.*

471 **Proof Sketch.** Given a timed automaton \mathcal{A} , we build a timed automaton \mathcal{A}^S that recognizes
 472 all suffixes of timed words in $\mathcal{L}(\mathcal{A})$. Given a fault model \mathcal{P} , we build an automaton $\mathcal{B}^{\mathcal{P}}$ from
 473 $\mathcal{A}_{\mathcal{P}}$ which remembers if a fault has occurred, and how many transitions were taken since a
 474 fault. Then, we re-label every transition occurring before a fault and till K steps after the

475 fault by ε , keeping the same locations, guards and resets, and leave transitions occurring
 476 more than K steps after a fault unchanged to obtain an automaton $\mathcal{B}^{\mathcal{P},\varepsilon}$. Accepting locations
 477 of $\mathcal{B}^{\mathcal{P}}$ are those of \mathcal{A} occurring after a fault in $\mathcal{B}^{\mathcal{P}}$. Then, every faulty run accepted by
 478 $\mathcal{B}^{\mathcal{P},\varepsilon}$ is associated with a word $\rho = (t_1, d_1) \dots (t_f, d_f)(t_{f+1}, d_{f+1}) \dots (t_{f+K}, d_{f+K}) \dots (t_n, d_n)$
 479 where t_1, \dots, t_{f+K} are ε transitions. A run ρ is BTN iff $(a_{f+K+1}, d_{f+K+1}) \dots (a_n, d_n)$ is a
 480 suffix of a timed word of \mathcal{A} , i.e., is recognized by \mathcal{A}^S . We can check that every word in
 481 $\mathcal{B}^{\mathcal{P},\varepsilon}$ (reading only ε before a fault) is recognized by the suffix automaton \mathcal{A}^S , by solving a
 482 language inclusion problem for timed automata with ε transitions. \blacktriangleleft

483 We note that ε -transitions are critical for the reduction of Proposition 19. To get
 484 decidability of K - \forall -resilience, it is thus necessary (but not sufficient) to be in a class with
 485 decidable timed language inclusion, such as Event-Recording timed automata [3], Integer
 486 Reset timed automata (IRTA) [18], or Strongly Non-Zeno timed automata [4]. However,
 487 to obtain decidability of K - \forall -resilience using Proposition 19, one needs also to ensure
 488 that inclusion is still decidable for automata in the presence of ε transitions. When a
 489 subclass C of timed automata is closed by enlargement (due to the fault model), and if timed
 490 language inclusion is decidable, even with ε transitions, then Proposition 19 implies that
 491 K - \forall -resilience is decidable for C . We show that this holds for the case of IRTA and leave
 492 other subclasses for future work. For IRTA [18], we know that $\mathcal{L}(\mathcal{A}) \subseteq \mathcal{L}(\mathcal{B})$ is decidable
 493 in EXPSPACE when \mathcal{B} is an IRTA [18] (even with ε transitions), from which we obtain an
 494 upper bound for K - \forall -resilience of IRTA. The enlargement of guards due to the fault can add
 495 transitions that reset clocks at non-integral times, but it turns out that the suffix automaton
 496 \mathcal{A}^S of Proposition 19 is still an IRTA. A matching lower bound is obtained by encoding
 497 inclusion for IRTA with K - \forall -resilience using a trick to replace the gadget in Proposition 18
 498 by an equivalent IRTA. Thus, we have Theorem 20 (proof in Appendix D).

499 \blacktriangleright **Theorem 20.** *K - \forall -resilience is EXPSPACE-Complete for IRTA.*

500 Finally, we conclude this section by remarking that universal *untimed* resilience is decidable
 501 for timed automata in general, using the reductions of Propositions 18 and 19:

502 \blacktriangleright **Theorem 21.** *Untimed K - \forall -resilience is EXPSPACE-Complete.*

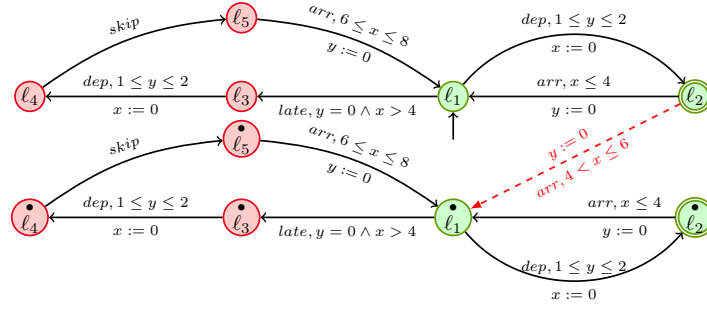
503 **Proof Sketch.** Untimed language inclusion of timed automata is EXPSPACE-Complete [9],
 504 so the reduction of Proposition 18 immediately gives the EXPSPACE lower bound. For
 505 the upper bound, we use the region construction and an ε -closure to build an automaton
 506 \mathcal{A}_U^S that recognizes untimed suffixes of words of \mathcal{A} , and an automaton $\mathcal{B}_U^{\mathcal{P}}$ that recognizes
 507 suffixes of words played K steps after a fault. Both are of exponential size. Then untimed
 508 K - \forall -resilience amounts to checking $\mathcal{L}(\mathcal{B}_U^{\mathcal{P}}) \subseteq \mathcal{L}(\mathcal{A}_U^S)$, yielding EXPSPACE upper bound. \blacktriangleleft

509 6 Conclusion

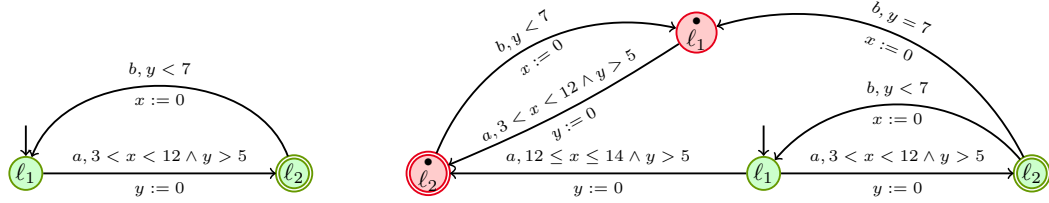
510 Resilience allows to check robustness of a timed system to unspecified delays. A universally
 511 resilient timed system recovers from any delay in some fixed number of steps. Existential
 512 resilience guarantees the existence of a controller that can bring back the system to a normal
 513 behavior within a fixed number of steps after an unexpected delay. Interestingly, we show
 514 that existential resilience enjoys better complexities/decidability than universal resilience.
 515 Universal resilience is decidable only for well behaved classes of timed automata such as IRTA,
 516 or in the untimed setting. A future work is to investigate resilience for other determinizable
 517 classes of timed automata, and a natural extension of resilience called *continuous resilience*,
 518 where a system recovers within some fixed duration rather than within some number of steps.

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■ **Figure 3** Faulty automata of the model of a train system with a mechanism to recover from delays as described in Example 8 and Figure 1



■ **Figure 4** \mathcal{A} on the left; Enlargement $\mathcal{A}_{\mathcal{P}}$ on the right, $\mathcal{P}(a) = 2, \mathcal{P}(b) = 0$.

569

A Example

570

► **Example 22.** Consider the automaton \mathcal{A} in Figure 4, with two locations ℓ_1 and ℓ_2 , a transition t_{12} from ℓ_1 to ℓ_2 and a transition t_{21} from ℓ_2 to ℓ_1 . The enlarged automaton $\mathcal{A}_{\mathcal{P}}$ has two extra locations ℓ_1, ℓ_2 , extra transitions between ℓ_1 and ℓ_2 , and from ℓ_1 to ℓ_2 and from ℓ_2 to ℓ_1 respectively. We represent a configuration of the automata with a pair $(\ell, \nu(x)|\nu(y))$ where, ℓ belongs to the set of the locations and $\nu(x)$ (resp. $\nu(y)$) represents the valuation of clock x (resp. clock y). Let $\rho_f = (\ell_1, 0|0) \xrightarrow{(t_{12},6)} (\ell_2, 6|0) \xrightarrow{(t_{21},13)} (\ell_1, 0|7) \xrightarrow{(t_{12},19)} (\ell_2, 4|0)$ be a *faulty run* reading the faulty word $(a, 6)(b, 13)(a, 19) \in \mathcal{L}(\mathcal{A}_{\mathcal{P}})$. This run is 1-BTN since the run $\sigma = (\ell_1, 0|0) \xrightarrow{(t_{12},6)} (\ell_2, 6|0) \xrightarrow{(t_{21},12)} (\ell_1, 0|6) \xrightarrow{(t_{12},19)} (\ell_2, 7|0)$ is an accepting run of \mathcal{A} , reading timed word $w_\sigma = (a, 6)(b, 12)(a, 19) \in \mathcal{L}(\mathcal{A})$. Similarly, the run $\rho' = (\ell_1, 0|0) \xrightarrow{(t_{12},14)} (\ell_2, 14|0) \xrightarrow{(t_{21},20)} (\ell_1, 0|6) \xrightarrow{(t_{12},31)} (\ell_2, 11|0)$ of $\mathcal{A}_{\mathcal{P}}$ reading word $(a, 14)(b, 20)(a, 31)$ is 1-BTN because of run $\sigma' = (\ell_1, 0|0) \xrightarrow{(t_{12},10)} (\ell_2, 10|0) \xrightarrow{(t_{21},15)} (\ell_1, 0|5) \xrightarrow{(t_{12},19)} (\ell_2, 4|0) \xrightarrow{(t_{21},20)} (\ell_1, 0|1) \xrightarrow{(t_{12},31)} (\ell_2, 11|0)$ reading the word $w_{\sigma'} = (a, 10)(b, 15)(a, 19)(b, 20)(a, 31)$. One can notice that ρ' and σ' are of different lengths. In fact, we can say something stronger, namely it is 1- \forall -resilient (and hence 1- \exists -resilient) as explained below.

584

The example consists of a single $(a.b)^*$ loop, where action a occurs between 3 and 12 time units after entering location ℓ_1 , and action b occurs less than 7 time units after entering ℓ_2 . A fault occurs either from ℓ_1 , in which case action a occurs $12 + d$ time units after entering ℓ_1 , with $d \in [0, 2]$, or from ℓ_2 , i.e., when b occurs exactly 7 time units after entering ℓ_2 . Once a fault has occurred, the iteration of a and b continues on ℓ_1 and ℓ_2 with non-faulty constraints. Consider a just faulty run ρ_f where fault occurs on event a . The timed word generated in ρ_f is of the form $w_f = (a, d_1).(b, d_2) \dots (a, d_k).(b, d_{k+1}).(a, d_{k+2})$, where $d_{k+2} = d_{k+1} + 12 + x$ with $x \in [0, 2]$. The word $w = (a, d_1).(b, d_2) \dots (a, d_k).(b, d_{k+1}).(a, d_{k+1} + 5).(b, d_{k+1} + 5 +$

591

592 $x).(a, d_{k+1} + 5 + x + 7)$ is also recognized by the normal automaton, and ends at date
 593 $d_{k+1} + 12 + x$. Hence, for every just faulty word w_f which delays action a , there exists a word
 594 w such for every timed word v , if $w_f.v$ is accepted by the faulty automaton, $w.v$ is accepted
 595 by the normal automaton. Now, consider a fault occurring when playing action b . The just
 596 faulty word ending with a fault is of the form $w_f = (a, d_1).(b, d_2) \dots (a, d_k).(b, d_k + 7)$. All
 597 occurrences of a occur at a date between $d_j + 3$ and $d_j + 12$ for some date d_j at which location ℓ_1
 598 is reached, (except the first time stamp $d_1 \in (5, 12)$) and all occurrences of b at a date strictly
 599 smaller than $d_i + 7$, where d_i is the date of last occurrence of a . Also, for any value $\epsilon \leq 7$ the
 600 word $w_\epsilon = (a, d_1).(b, d_2) \dots (a, d_k).(b, d_k + 7 - \epsilon)$ is non-faulty. Let $v_1 = 12 - d_1$, recall that
 601 $d_1 \in (5, 12)$. If we choose $\epsilon < v_1$ then the run $w_\epsilon^+ = (a, d_1 + \epsilon).(b, d_2 + \epsilon) \dots (a, d_k + \epsilon).(b, d_k + 7)$
 602 is also non-faulty because $5 < d_1 + \epsilon < d_1 + v_1 = 12$. Clearly, we can extend w_ϵ^+ to match
 603 transitions fired from w_ϵ hence, the automaton of the example is 1- \forall -resilient.

604 B Proofs for section 4

605 **► Lemma 12** *There exists $\nu_2 \models r$ and a timed run ρ'_2 of \mathcal{A} with the same duration as ρ_2 ,*
 606 *such that $\rho_2 \otimes \rho'_2$ ends in state $(l_{\mathcal{A}\mathcal{P}}, l_{\mathcal{A}}, K, \nu_2)$.*

607 **Proof.** Let t_1, \dots, t_n be the sequence of transitions of ρ_1, ρ_2 taken respectively, at dates
 608 d_1, \dots, d_n and e_1, \dots, e_n . Similarly, we will denote by t'_1, \dots, t'_k the sequence of transitions
 609 of ρ'_1 , taken at dates d'_1, \dots, d'_k . Run ρ'_2 will pass by the same transitions t'_1, \dots, t'_k , but with
 610 possibly different dates e'_1, \dots, e'_k such that:

- 611 ■ the duration of ρ'_2 is the same as the duration of ρ_2 ,
- 612 ■ $\tilde{\sigma}_{\rho'_2}$ follows the same sequence of states of $\mathcal{R}_{\text{strong}}(\mathcal{A})$ as $\tilde{\sigma}_{\rho'_1}$ (in particular, ρ'_2 is a valid
 613 run as it fulfils the guards of its transitions, which are the same as those of ρ'_1).
- 614 ■ $\tilde{\sigma}_{\rho_2 \otimes \rho'_2}$ reaches the same state of $\mathcal{R}_{\text{strong}}(\mathcal{A}\mathcal{P} \otimes_K \mathcal{A})$ as $\tilde{\sigma}_{\rho_1 \otimes \rho'_1}$.

615 We translate these into three successive requirements on the dates $(e'_i)_{i \leq k}$ of ρ'_2 :

- 616 R1. The first requirement is $e'_k = e_n$,
- 617 R2. The second requirement sets the integral part $\lfloor e'_i \rfloor = \lfloor d'_i \rfloor$ for all $i \leq k$. Remark that we
 618 already have $\lfloor e'_k \rfloor = \lfloor e_n \rfloor = \lfloor d_n \rfloor = \lfloor d'_k \rfloor$ by the first requirement and the hypothesis,
- 619 R3. The third requirement tackles the fractional part $(\text{frac}(e'_i))_{i \leq k}$. It is given as a set
 620 of satisfiable constraints, defined hereafter as a partial ordering on $(\text{frac}(e'_i))_{i \leq k} \cup$
 621 $(\text{frac}(e_i))_{i \leq n}$.

622 Notice that the value of a clock x just before firing transition t_i is obtained by considering
 623 the date d_i of t_i minus the date d_i^x of the latest transition $t_j, j < i$ at which x has been
 624 last reset before i . In particular, the difference $x - y$ between clocks x, y just before firing
 625 transition t_i is $(d_i - d_i^x) - (d_i - d_i^y) = d_i^y - d_i^x$. That is, the value of a clock or its difference
 626 can be obtained by considering the difference between two dates of transitions. A constraint c
 627 given by $x - y \in (n, n + 1)$ is equivalent with the constraint $d(c)$ given by $d_i^y - d_i^x \in (n, n + 1)$.

628 We then characterize the conditions required for the run $\rho_2 \otimes \rho'_2$ to reach the same region
 629 r of $\mathcal{R}_{\text{strong}}(\mathcal{A}\mathcal{P} \otimes_K \mathcal{A})$ which was reached by $\rho_1 \otimes \rho'_1$. These conditions are described as on
 630 region r in the following equivalent ways:

- 631 1. A set of constraints C on the disjoint union $X'' = X_{\mathcal{A}\mathcal{P}} \uplus X_{\mathcal{A}}$ of clocks of $\mathcal{A}\mathcal{P}$ and \mathcal{A} , of
 632 the form $x - y \in (n, n + 1)$ or $x - y = n$ or $x - y > \text{Max}$ (possibly considering a null
 633 clock y) for $n \in \mathbb{Z}$,

- 634 2. The associated set of constraints $C' = \{d(c) \mid c \in C\}$ on $D = \{d_x \mid x \in X_{\mathcal{A}_P}\} \uplus \{d_{x'} \mid$
635 $x' \in X_{\mathcal{A}}\}$, with d_x the date of the latest transition t_j^{\otimes} that resets the clock $x \in X_{\mathcal{A}_P}$, and
636 $d_{x'}$ the date of the latest transition t_i^{\otimes} that resets clock $x' \in X_{\mathcal{A}}$,
- 637 3. An ordering \leq' over $FP = \{\text{frac}(\tau) \mid \tau \in D\}$ defined as follows: for each constraint
638 $\tau - \tau' \in (n, n+1)$ of C' , if $\lfloor \tau \rfloor = \lfloor \tau' \rfloor + n$ then $\text{frac}(\tau) <' \text{frac}(\tau')$, and if $\lfloor \tau \rfloor = \lfloor \tau' \rfloor + n + 1$
639 then $\text{frac}(\tau') <' \text{frac}(\tau)$.
- 640 For each constraint $\tau - \tau' = n$ of C' , then $\text{frac}(\tau') =' \text{frac}(\tau)$.
- 641 For each constraint $\tau - \tau' > c_{\max}$ of C' such that $\lfloor \tau \rfloor = \lfloor \tau' \rfloor + c_{\max}$, we have $\text{frac}(\tau') >'$
642 $\text{frac}(\tau)$ (if $\lfloor \tau \rfloor \geq \lfloor \tau' \rfloor + c_{\max} + 1$, then we dont need to do anything), where $c_{\max} =$
643 $\max(\{c_x \mid x \in X\})$.

644 Further, path ρ'_2 needs to visit the regions r_1, \dots, r_k visited by ρ'_1 . For each i , visiting
645 region r_i is characterized by a set of constraints C_i , which we translate as above as an
646 ordering \leq'_i on $FP' = \{\text{frac}(d'_i) \mid i \leq k\}$.

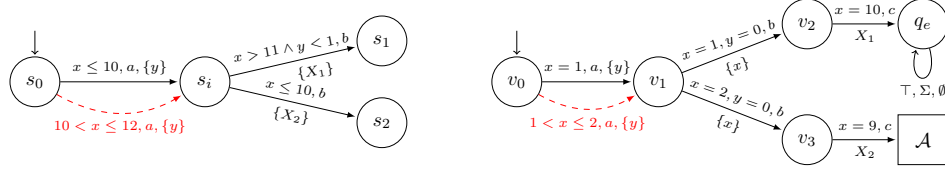
647 Thus, finally, we can collect all the requirements for having ρ' with required properties by
648 defining \leq'' over $FP' \cup FP$ (notice that it is not a disjoint union) as the transitive closure of
649 the union of all \leq'_i and of \leq' . As the union of constraints on C'_i and on C' is satisfied by the
650 dates $(d_i)_{i \leq n}$ and $(d'_i)_{i \leq k}$ of ρ_1 and ρ'_1 , the union of constraints is satisfiable. Equivalently,
651 \leq'' is a partial ordering, respecting the total natural ordering \leq on $FP \cup FP'$. We will
652 denote $\tau ='' \tau'$ whenever $\tau \leq'' \tau'$ and $\tau' \leq'' \tau$, and $\tau <'' \tau'$ if $\tau \leq'' \tau'$ but we dont have
653 $\tau ='' \tau'$. Because \leq'' is a partial ordering, there is no τ, τ' with $\tau <'' \tau' <'' \tau$.

654 Note that there is only one way of fulfilling the first two requirements R1. and R2; namely
655 by matching e'_k and e_n , and by witnessing dates with the same integral parts in e'_k, e_n as
656 well as d'_k, d_n . While this takes care of the last values, to obtain the remaining values, we
657 can apply any greedy algorithm fixing successively $\text{frac}(e'_{k-1}) \dots \text{frac}(e'_1)$ and respecting \leq''
658 to yield the desired result. We provide a concrete such algorithm for completeness:

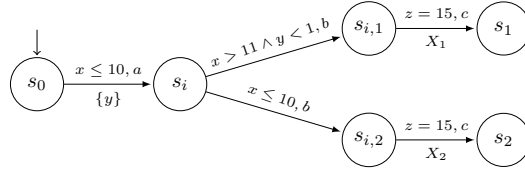
659 We will start from the fixed value of $\text{frac}(e'_{k-1})$ and work backwards. Let us assume
660 inductively that $\text{frac}(e'_{k-1}) \dots \text{frac}(e'_{i+1})$ have been fixed. We now describe how to obtain
661 $\text{frac}(e'_i)$. If $\text{frac}(d'_i) ='' \text{frac}(d'_j)$, $j > i$ then we set $\text{frac}(e'_i) = \text{frac}(e'_j)$. If $\text{frac}(d'_i) ='' \text{frac}(d_j)$,
662 then we set $\text{frac}(e'_i) = \text{frac}(e_j)$. Otherwise, consider the sets $L_i = \{\text{frac}(e_j) \mid j \leq n, \text{frac}(d_j) <''$
663 $\text{frac}(d'_i)\} \cup \{\text{frac}(e'_j) \mid i < j \leq n, \text{frac}(d'_j) <'' \text{frac}(d'_i)\}$. Also, consider $U_i = \{\text{frac}(e_j) \mid j \leq$
664 $n, \text{frac}(d_j) >'' \text{frac}(d'_i)\} \cup \{\text{frac}(e'_j) \mid i < j \leq n, \text{frac}(d'_j) >'' \text{frac}(d'_i)\}$. We let $l_i = \max(L_i)$
665 and $u_i = \min(U_i)$. We then set $\text{frac}(e'_i)$ to any value in (l_i, u_i) . It remains to show that we
666 always have $l_i < u_i$, which will show that such a choice of value for the fractional part of e'_i
667 is indeed possible.

668 By contradiction, consider that there exists i such that $l_i \geq u_i$, and consider the
669 maximal (first) such i . First, assume that both l_i and u_i are of the form $\text{frac}(e_j), \text{frac}(e_k)$
670 respectively, i.e. corresponds to clock values in the last regions of ρ_2 . The contradiction
671 hypothesis is $l_i = \text{frac}(e_j) \geq u_i = \text{frac}(e_k)$. By definition of L_i and U_i , we also have
672 $\text{frac}(d_j) <'' \text{frac}(d'_i) <'' \text{frac}(d_k)$. In particular, $\text{frac}(d_j) < \text{frac}(d_k)$. This is a contradiction
673 with $\tilde{\sigma}_{\rho_1} = \tilde{\sigma}_{\rho_2}$, as the strong region reached by ρ_1 and ρ_2 are the same. A contradiction.

674 Otherwise, at least one of l_i, u_i is of the form $\text{frac}(e'_j)$, with $j > i$ (consider j minimal
675 if both are of this form). By symetry, let say $l_i = \text{frac}(e'_j) \geq u_i$. Let say $u_i = \text{frac}(e_k)$,
676 as $u_i = \text{frac}(e'_k)$ with $k > j$ is similar since it has been fixed before $\text{frac}(e'_j)$. We have
677 $\text{frac}(d'_j) <'' d'_i <'' \text{frac}(d_k)$ by definition of L_i, U_i . In particular $\text{frac}(d'_j) <'' \text{frac}(d_k)$: That
678 is, $k \in U_j$, and by construction, and as $j > i$, we have $l_i = \text{frac}(e'_j) < \text{frac}(e_k) = u_i$, a
679 contradiction. \blacktriangleleft



■ **Figure 5** The gadgets \mathcal{G} (left) and $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ (right) which is untimed 2- \exists -resilient iff $\mathcal{L}(\mathcal{A}) \neq \emptyset$.



■ **Figure 6** The gadget automaton \mathcal{G}_{und} .

B.1 Hardness for K - \exists -resilience

680

► **Theorem 15** *The K - \exists -resilience problem for timed automata is PSPACE-Hard.*

681

682 **Proof.** We proceed by reduction from the language emptiness problem, which is known
 683 to be PSPACE-Complete for timed automata. We can reuse the gadget \mathcal{G}_{und} of Figure 6.
 684 We take any automaton \mathcal{A} and collapse its initial state to state s_1 in the gadget. We
 685 recall that s_1 is accessible at date 15 only after a fault. We add a self loop with transition,
 686 $t_e = (s_2, \sigma, true, \emptyset, s_2)$ for every $\sigma \in \Sigma$. This means that after reaching s_2 , which is accessible
 687 only at date 15 if no fault has occurred, the automaton accepts any letter with any timing.
 688 Then, if \mathcal{A} has no accepting word, there is no timed word after a fault which is a suffix
 689 of a word in $\mathcal{L}(\mathcal{A})$, and conversely, if $\mathcal{L}(\mathcal{A}) \neq \emptyset$, then any word recognized from s_1 is
 690 also recognized from q_e . So the language emptiness problem reduces to a 2- \exists -resilience
 691 question. ◀

B.2 Untimed K - \exists -resilience

692

► **Theorem 16** *Untimed K - \exists -resilience is PSPACE-Complete.*

693

694 **Proof. Membership :** For every run of \mathcal{A} , there is a path in $\mathcal{R}(\mathcal{A})$. So, \mathcal{A} is untimed
 695 K - \exists -resilient if and only if, for all states q reached by a just faulty run, there exists a
 696 maximal accepting path σ from q such that, K steps after, the sequence of actions on its
 697 suffix σ_s agrees with that of an accepting path σ in $\mathcal{R}(\mathcal{A})$. We now prove that this property
 698 can be verified in PSPACE.

699 Let $q = (l, r)$ be a state of $\mathcal{R}(\mathcal{A}_{\mathcal{P}})$ reached after a just faulty run. K steps after reaching
 700 $q = (l, r)$ of $\mathcal{R}(\mathcal{A}_{\mathcal{P}})$, one can check in PSPACE, if there exists a path σ_s whose sequence
 701 of actions is the same as the suffix of an accepting path σ of $\mathcal{R}(\mathcal{A})$. That is, either both
 702 these end in a pair of accepting states from which no transitions are defined (both paths are
 703 maximal), or visit a pair of states twice such that the cyclic part of the path contains both
 704 an accepting state of $\mathcal{R}(\mathcal{A}_{\mathcal{P}})$ and an accepting state of $\mathcal{R}(\mathcal{A})$. To find these paths σ, σ_s , one
 705 just needs to guess them, i.e., build them synchronously by adding a pair of transitions to
 706 the already built path only if they have the same label. One needs to remember the current
 707 pair of states reached, and possibly guess a pair of states $(s_{\mathcal{A}}, s_{\mathcal{A}_{\mathcal{P}}})$ on which a cycle starts,
 708 and two bits $b_{\mathcal{A}}$ (resp. $b_{\mathcal{A}_{\mathcal{P}}}$) to remember if an accepting state of \mathcal{A} (resp. $\mathcal{A}_{\mathcal{P}}$) has been seen
 709 since $(s_{\mathcal{A}}, s_{\mathcal{A}_{\mathcal{P}}})$. A maximal finite path or a lasso can be found on a path of length smaller
 710 than $|\mathcal{R}(\mathcal{A}_{\mathcal{P}})| \times |\mathcal{R}(\mathcal{A})|$, and the size of the currently explored path can be memorized with

711 $\log_2(|\mathcal{R}(\mathcal{A}_{\mathcal{P}})| \times |\mathcal{R}(\mathcal{A})|)$ bits. This can be done in PSPACE. The complement of this, i.e.,
 712 checking that no maximal path originating from q with the same labeling as a suffix of a
 713 word recognized by $\mathcal{R}(\mathcal{A})$ K steps after a fault exists, is in PSPACE too.

714 Now, to show that \mathcal{A} is *not* untimed K - \exists -resilient, we simply have to find one untimed
 715 non- K - \exists -resilient witness state q reachable immediately after a fault. To find it, non
 716 deterministically guess such a witness state q along with a path of length not more than the
 717 size of $|\mathcal{R}(\mathcal{A}_{\mathcal{P}})|$ and apply the PSPACE procedure above to decide whether it is a untimed
 718 non- K - \exists -resilience witness. Guess of q is non-deterministic, which gives an overall NPSpace
 719 complexity, but again, using Savitch's theorem, we can say that untimed K - \exists -resilience is
 720 in PSPACE.

721 *Hardness* : We can now show that untimed K - \exists -resilience is PSPACE-Hard. Consider a
 722 timed automaton \mathcal{A} with alphabet Σ and the construction of an automata that uses a gadget
 723 shown in Figure 5 (right). Let us call this automaton $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$. This automaton reads a word
 724 $(a, 1).(b, 1).(c, 11)$ and then accepts all timed words 2 steps after a fault, via Σ loop on a
 725 particular accepting state q_e . If $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ takes the faulty transition (marked in dotted red)
 726 then it resets all clocks of \mathcal{A} and behaves as \mathcal{A} . The accepting states are $q_e \cup F$. Then, \mathcal{A}
 727 has an accepting word if and only if $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ is untimed 2- \exists -resilient. Since the emptiness
 728 problem for timed automata is PSPACE-Complete, the result follows. ◀

729 C Proofs for section 5

730 ▶ **Proposition 18** *Language inclusion for timed automata can be reduced to K - \forall -resilience.*
 731 *Thus, K - \forall -resilience is undecidable in general for timed automata.*

732 **Proof.** Let $\mathcal{A}_1 = (L_1, \{l_{0_1}\}, X_1, \Sigma_1, T_1, F_1)$ and $\mathcal{A}_2 = (L_2, \{l_{0_2}\}, X_2, \Sigma_2, T_2, F_2)$ be two timed
 733 automata with only one initial state (w.l.o.g). We build a timed automaton \mathcal{B} such that
 734 $\mathcal{L}(\mathcal{A}_1) \subseteq \mathcal{L}(\mathcal{A}_2)$ if and only if \mathcal{B} is 2- \forall -resilient.

735 We first define a gadget \mathcal{G}_{und} that allows to reach a state s_1 at an arbitrary date $d_1 = 15$
 736 when a fault happens, and a state s_2 at date $d_2 = d_1 = 15$ when no fault occur. This gadget
 737 is shown in Fig 6. \mathcal{G}_{und} has 6 locations $s_0, s_i, s_{i,1}, s_1, s_2 \notin L_1 \cup L_2$, three new clocks $x, y, z \notin$
 738 $X_1 \cup X_2$, three new actions $a, b, c \notin \Sigma_1 \cup \Sigma_2$, and 5 transitions $t_0, t_1, t_2, t_3, t_4 \notin T_1 \cup T_2$ defined
 739 as: $t_0 = (s_0, a, g_0, \{y\}, s_i)$ with $g_0 ::= x \leq 10$, $t_1 = (s_i, b, g_1, \emptyset, s_{i,1})$ with $g_1 ::= x > 11 \wedge y < 1$,
 740 $t_2 = (s_i, b, g_2, \emptyset, s_{i,2})$ with $g_2 ::= x \leq 10$, $t_3 = (s_{i,1}, c, g_3, X_1, s_1)$ with $g_3 ::= z = 15$, and
 741 $t_4 = (s_{i,2}, c, g_4, X_2, s_2)$ with $g_4 ::= z = 15$. Clearly, in this gadget, transition t_1 can never
 742 fire, as a configuration with $x > 11$ and $y < 1$ is not accessible.

743 We build a timed automaton \mathcal{B} that contains all transitions of \mathcal{A}_1 and \mathcal{A}_2 , but preceded
 744 by \mathcal{G}_{und} by collapsing the initial location of \mathcal{A}_1 i.e., l_{0_1} with s_1 and the initial location of \mathcal{A}_2
 745 i.e., l_{0_2} with s_2 . We also use a fault model $\mathcal{P} : a \rightarrow [0, 2]$, that can delay transitions t_0 with
 746 action a by up to 2 time units. The language $\mathcal{L}(\mathcal{B})$ is the set of words:

$$747 \begin{aligned} \mathcal{L}(\mathcal{B}) = \{ & (a, d_1)(b, d_2)(c, 15)(\sigma_1, d_3) \dots (\sigma_n, d_{n+2}) \mid (d_1 \leq 10) \\ & \wedge (d_2 \leq 10) \wedge (d_2 - d_1 < 1) \\ & \wedge \exists w = (\sigma_1, d'_3) \dots (\sigma_n, d'_{n+2}) \in \mathcal{L}(\mathcal{A}_2), \forall i \in 3..n + 2, d'_i = d_i - 15 \} \end{aligned}$$

748 The enlargement of \mathcal{B} is denoted by $\mathcal{B}_{\mathcal{P}}$. The words in $\mathcal{L}(\mathcal{B}_{\mathcal{P}})$ is the set of words in $\mathcal{L}(\mathcal{B})$
 749 (when there is no fault) plus the set of words in:

$$750 \begin{aligned} \mathcal{L}^F(\mathcal{B}_{\mathcal{P}}) \{ & (a, d_1)(b, d_2)(c, 15)(\sigma_1, d_3) \dots (\sigma_n, d_{n+2}) \mid (10 < d_1 \leq 12) \\ & \wedge d_2 > 11 \wedge (d_2 - d_1 < 1) \\ & \wedge \exists w = (\sigma_1, d'_3) \dots (\sigma_n, d'_{n+2}) \in \mathcal{L}(\mathcal{A}_1), \forall i \in 3..n + 2, d'_i = d_i - 15 \} \end{aligned}$$

751 Now, \mathcal{B} is K - \forall -resilient for $K = 2$ if and only if every word in $\mathcal{L}^F(\mathcal{B}_{\mathcal{P}})$ is BTN after 2
 752 steps ($K = 2$), i.e., for every word $w = (a, d_1)(b, d_2)(c, 15)(\sigma_1, d_3) \dots (\sigma_n, d_{n+2})$ in $\mathcal{L}^F(\mathcal{B}_{\mathcal{P}})$,

753 if there exists a word $w = (a, d'_1)(b, d'_2)(c, 15)(\sigma_1, d_3) \dots (\sigma_n, d_{n+2})$ in $\mathcal{L}(\mathcal{B})$. This means that
 754 every word of \mathcal{A}_1 is a word of \mathcal{A}_2 . So $\mathcal{L}(\mathcal{A}_1) \subseteq \mathcal{L}(\mathcal{A}_2)$ if and only if \mathcal{B} is 2- \forall -resilient.

755 As language inclusion for timed automata is undecidable [2], an immediate consequence
 756 is that K - \forall -resilience of timed automata is undecidable.

757 ◀

758 **► Proposition 19** K - \forall -resilience can be reduced to language inclusion for timed automata
 759 with ε -transitions.

760 **Proof.** Given a timed automaton $\mathcal{A} = (L, I, X, \Sigma, T, F)$, we can build a timed automaton
 761 \mathcal{A}^S that recognizes all suffixes of timed words recognized by \mathcal{A} . Formally, \mathcal{A}^S contains the
 762 original locations and transitions of \mathcal{A} , a copy of all location, a copy of all transitions where
 763 letters are replaced by ε , and a transition from copies to original locations labeled by their
 764 original letters.

765 We have $\mathcal{A}^S = (L^S, I^S, X, \Sigma \cup \{\varepsilon\}, T^S, F)$, where $L^S = L \cup \{l' \mid l \in L\}$, $I^S = \{l' \in L_S, l \in$
 766 $I\} \cup T^S = T \cup \{(l'_1, g, \varepsilon, R, l'_2) \mid \exists (l_1, g, \sigma, R, l_2) \in T\} \cup \{(l'_1, g, \sigma, R, l_2) \mid \exists (l_1, g, \sigma, R, l_2) \in T\}$.
 767 Obviously, for every timed word $(a_1, d_1)(a_2, d_2) \dots (a_n, d_n)$ recognized by \mathcal{A} , and every
 768 index $k \in 1..n$, the words $(\varepsilon, d_1)(\varepsilon, d_k)(a_{k+1}, d_{k+1}) \dots (a_n, d_n) = (a_{k+1}, d_{k+1}) \dots (a_n, d_n)$ is
 769 recognized by \mathcal{A}^S .

770 Given a timed automaton \mathcal{A} and a fault model \mathcal{P} , we build an automaton $\mathcal{B}^{\mathcal{P}}$ which
 771 remembers if a fault has occurred, and how many transitions have been taken since a fault.

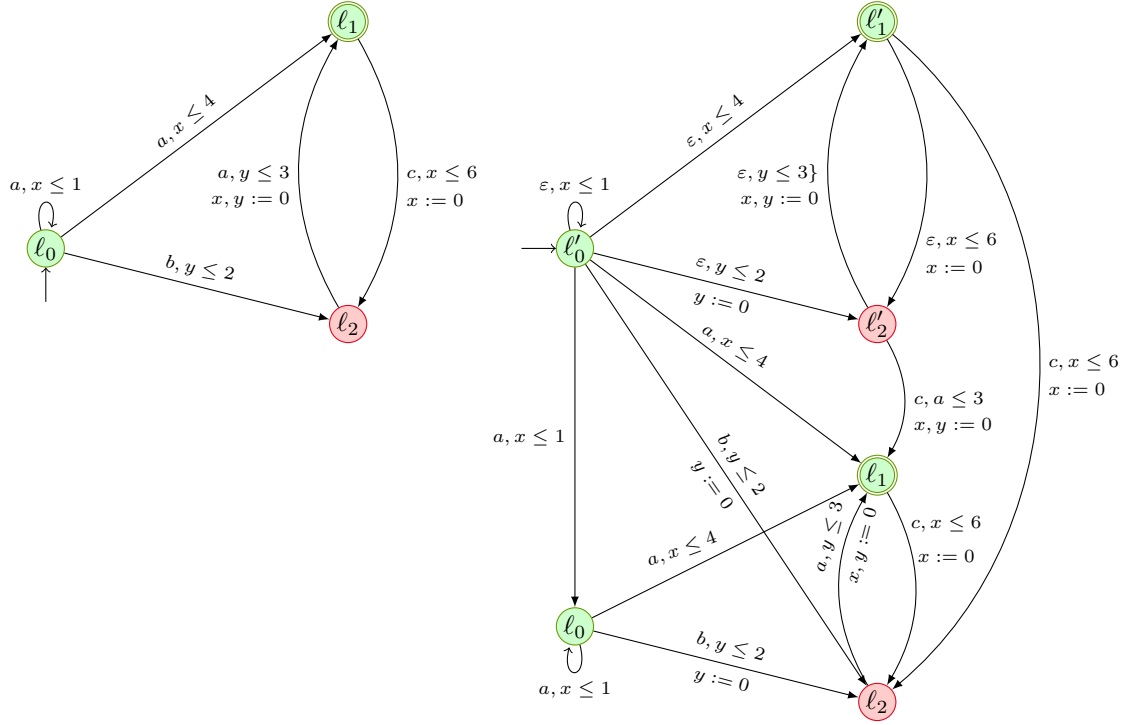
772 **► Definition 23** (Counting automaton). Let $\mathcal{A}_{\mathcal{P}} = (L, I, X, \Sigma, T, F)$ and be a timed automaton
 773 with faulty transitions. Let $K \in \mathbb{N}$ be an integer. Then, the faulty automaton $\mathcal{B}^{\mathcal{P}}$ is a tuple
 774 $\mathcal{B}^{\mathcal{P}} = (L^{\mathcal{P}}, I^{\mathcal{P}}, X, \Sigma, T^{\mathcal{P}}, F^{\mathcal{P}})$ where $L^{\mathcal{P}} \subseteq \{L \times \{0\}\}$, $F^{\mathcal{P}} = F \times [-1, K]$, and initial set of
 775 states $I^{\mathcal{P}} = I \times \{-1\}$. Intuitively, -1 means no fault has occurred yet. Then we assign K
 776 and decrement to 0 to denote that K steps after fault have passed. The set of transitions $T^{\mathcal{P}}$
 777 is as follows: We have $((l, n), g, a, R, (l', n')) \in T^{\mathcal{P}}$ if and only if either:

- 778 \blacksquare $n \neq 0$ (no fault has occurred, or less than K steps of \mathcal{B} have occurred), we have transition
 779 $t = (l, g, a, R, l) \in T$, and either: $n = -1$, the transition t is faulty and $n' = K$, or
 780 $n = -1$, the transition t is non faulty and $n' = -1$, or $n > 0$ and $n' = n - 1$.
- 781 \blacksquare $n = n' = 0$ (at least K steps after a fault have occurred), and there exists a transition
 782 $t = (l, g, a, R, l') \in T$.

783 Then, we can build an automaton $\mathcal{B}^{\mathcal{P}, \varepsilon}$ by re-labeling every transition occurring before
 784 a fault and until K steps after the fault by ε , keeping the same locations, guards and
 785 resets, and leave transitions occurring more than K steps after a fault unchanged. The
 786 relabeled transitions are transitions starting from a location (l, n) with $n \neq 0$. Accepting
 787 locations of $\mathcal{B}^{\mathcal{P}, \varepsilon}$ are of the form $(l, 0)$ where l is an accepting locations of \mathcal{A} occurring after
 788 a fault in $\mathcal{B}^{\mathcal{P}}$. Then, every faulty run accepted by $\mathcal{B}^{\mathcal{P}, \varepsilon}$ is associated with a word of the
 789 form $\rho = (t_1, d_1) \dots (t_f, d_f)(t_{f+1}, d_{f+1}) \dots (t_{f+K}, d_{f+K}) \dots (t_n, d_n)$ where t_1, \dots, t_{f+K} are ε
 790 transitions. A run ρ is BTN if and only if $(a_{f+K+1}, d_{f+K+1}) \dots (a_n, d_n)$ is a suffix of a timed
 791 word of \mathcal{A} , i.e., is recognized by \mathcal{A}^S .

792 Now one can check that every word in $\mathcal{B}^{\mathcal{P}, \varepsilon}$ (reading only ε before that fault) is recognized
 793 by the suffix automaton \mathcal{A}^S , i.e. solve a language inclusion problem for timed automata with
 794 ε transitions. ◀

795 **► Theorem 21** Untimed K - \forall -resilience is EXPSPACE-Complete.



■ **Figure 7** An example automaton \mathcal{A} (left) and its the suffix automaton \mathcal{A}^S (right)

796 **Proof.** Recall that untimed language inclusion of timed automata is EXPSPACE-Complete [9].
 797 The lower bound is readily obtained by using the reduction of Proposition 18.

798 For the upper bound, we will use the construction of automata \mathcal{A}^S and $\mathcal{B}^{\mathcal{P}, \varepsilon}$ built during
 799 the reduction of Proposition 19. We however need inclusion of TA with ε transitions, and
 800 thus we adapt the EXPSPACE algorithm in the presence of ε transitions:

801 We can consider ε transitions as transitions labeled by any letter, and build the region
 802 automata $\mathcal{A}_{\#} = \mathcal{R}(\mathcal{A}^S)$ and $\mathcal{B}_{\#} = \mathcal{R}(\mathcal{B}^{\mathcal{P}, \varepsilon})$. These automata are untimed automata of size
 803 exponential in the number of clocks, with ε transitions. We can perform an ε reduction
 804 on $\mathcal{A}_{\#}$ to obtain an automaton $\mathcal{A}_{\#}^S$ with the same number of states as $\mathcal{A}_{\#}$ that recognizes
 805 untimed suffixes of words of \mathcal{A} . Similarly, we can perform an ε reduction on $\mathcal{B}_{\#}$ to obtain an
 806 automaton $\mathcal{B}_{\#}^{\mathcal{P}}$ with the same number of states as $\mathcal{B}_{\#}$ that recognizes suffixes of words played
 807 K steps after a fault.

808 We then use a usual PSPACE inclusion algorithm to check that $\mathcal{L}(\mathcal{B}_{\#}^{\mathcal{P}}) \subseteq \mathcal{L}(\mathcal{A}_{\#}^S)$, which
 809 yields the EXPSPACE upper bound, as $\mathcal{A}_{\#}^S, \mathcal{B}_{\#}^{\mathcal{P}}$ have an exponential number of states w.r.t.
 810 $|\mathcal{A}|$. ◀

811 **D Resilience of Integer Reset Timed Automata**

812 Let us recall some elements used to prove decidability of language inclusion in IRTA. For
 813 a given IRTA \mathcal{A} we can define a map $f : \rho \rightarrow w_{unt}$ that maps every run ρ of \mathcal{A} to an
 814 untimed word $w_{unt} \in (\{\checkmark, \delta\} \cup \Sigma)^*$. For a real number x with $k = \lfloor x \rfloor$, we define a map
 815 $dt(x)$ from \mathbb{R} to $\{\checkmark, \delta\}^*$ as follows : $dt(x) = (\delta.\checkmark)^k$ if x is integral, and $dt(x) = (\delta.\checkmark)^k.\delta$
 816 otherwise. Then, for two reals $x < y$, the map $dte(x, y)$ is the suffix that is added to $dt(x)$
 817 to obtain $dt(y)$. Last, the map f associates to a word $w = (a_1, d_1) \dots (a_n, d_n)$ the word

818 $f(w) = w_1.a_1.w_2.a_2 \dots w_n.a_n$ where each w_i is the word $w_i = dte(d_{i-1}, d_i)$. The map f maps
 819 global time elapse to a word of \checkmark and δ but keeps actions unchanged. We define another map
 820 $f_{\downarrow} : w \rightarrow \{\checkmark, \delta\}^*$ that maps every word w of \mathcal{A} to a word in $\{\checkmark, \delta\}^*$ dropping the actions from
 821 $f(w)$. Consider for example, a word $w = (a, 1.6)(b, 2.7)(c, 3.4)$ then, $f(w) = \delta\checkmark\delta a\checkmark\delta b\checkmark\delta c$,
 822 and $f_{\downarrow}(w) = \delta\checkmark\delta\checkmark\delta\checkmark\delta$. It is shown in [18] for two timed words ρ_1, ρ_2 with $f(\rho_1) = f(\rho_2)$
 823 then $\rho_1 \in \mathcal{L}(\mathcal{A})$ if and only if $\rho_2 \in \mathcal{L}(\mathcal{A})$. It is also shown that we can construct a Marked
 824 Timed Automata (MA) from \mathcal{A} with one extra clock and polynomial increase in the number
 825 of locations such that $Unt(\mathcal{L}(MA)) = f(\mathcal{L}(\mathcal{A}))$. The MA of \mathcal{A} duplicates transitions of \mathcal{A} to
 826 differentiate firing at integral/non integral dates, plus transitions that make time elapsing
 827 visible using the additional clock which is reset at each global integral time stamp.

828 **► Definition 24** (Marked Timed Automaton (MA)). *Given a timed automaton $\mathcal{A} = (L, L_0, X, \Sigma, T, F)$
 829 the Marked Timed Automata of \mathcal{A} is a tuple $MA = (L', L'_0, X \cup \{n\}, \Sigma \cup \{\checkmark, \delta\}, T', F')$ such
 830 that*

- 831 i) $n \notin X$
- 832 ii) $L' = L^0 \cup L^+$ where for $\alpha \in \{0, +\}$, $L^\alpha = \{l^\alpha \mid l \in L\}$
- 833 iii) $L'_0 = \{l^0 \mid l \in L_0\}$,
- 834 iv) $F' = \{l^0, l^+ \mid l \in F\}$ and
- 835 v) T' is defined as follows,

$$836 \quad T' = \{(l^0, a, g \wedge n = 0?, R, l^0) \mid (l, a, g, R, l') \in E\}$$

$$837 \quad \cup \{(l^+, a, g \wedge 0 < n < 1?, R, l^+) \mid (l, a, g, R, l') \in E\}$$

$$838 \quad \cup \bigcup_{l \in L} \{(l^0, \delta, 0 < n < 1, \emptyset, l^+)\} \cup \bigcup_{l \in L} \{(l^+, \checkmark, n = 1?, \{n\}, l^0)\}$$

$$839$$

840 Then we have the following results.

841 **► Theorem 25** ([18]Thm.5). *Let \mathcal{A} be a timed automaton and MA be its marked automaton.
 842 Then $Unt(\mathcal{L}(MA)) = f(\mathcal{L}(\mathcal{A}))$*

843 **► Remark 26.** The marked timed automaton of an IRTA is also an IRTA.

844 The proofs of resilience for IRTA will also rely on the following properties,

845 **► Theorem 27** (Thm.3, [18]). *If \mathcal{A} is an IRTA and $f(w) = f(w')$, then $w \in \mathcal{L}(\mathcal{A})$ if and
 846 only if $w' \in \mathcal{L}(\mathcal{A})$*

847 **► Lemma 28.** *The timed suffix language of an IRTA \mathcal{A} can be recognized by an ε -IRTA \mathcal{A}^S*

848 **Proof.** Let $\mathcal{A} = (L, X, \Sigma, T, \mathcal{G}, F)$ be a timed automaton. We create an automaton $\mathcal{A}^S =$
 849 $(L^S, X, \Sigma \cup \{\varepsilon\}, T^S, \mathcal{G}, F)$ as follows. We set $L^S = L \cup L_\varepsilon$, where $L_\varepsilon = \{l_\varepsilon \mid l \in L\}$ i.e., L^S
 850 contains a copy of locations in \mathcal{A} and another ‘‘silent’’ copy. The initial location of \mathcal{A}^S is $l_{0,\varepsilon}$.
 851 We set $T^S = T \cup T_\varepsilon \cup T'_\varepsilon$, where $T_\varepsilon = \{(l_\varepsilon, \varepsilon, true, \emptyset, l) \mid l \in L\}$ and $T'_\varepsilon = \{(l_\varepsilon, \varepsilon, g, R, l'_\varepsilon) \mid$
 852 $\exists (l, a, g, R, l') \in T\}$. Clearly, for every timed word $w = (a_1, d_1) \dots (a_i, d_i)(a_{i+1}, d_{i+1}) \dots (a_n, d_n)$
 853 of $\mathcal{L}(\mathcal{A})$ and index i , the word $w' = (\varepsilon, d_1) \dots (\varepsilon, d_i)(a_{i+1}, d_{i+1}) \dots (a_n, d_n) = (a_{i+1}, d_{i+1}) \dots (a_n, d_n)$
 854 is a recognized by \mathcal{A}^S , and it is easy to verify that \mathcal{A}^S is an ε -IRTA. ◀

855 **► Lemma 29.** *For two IRTA \mathcal{A} and \mathcal{B} and their corresponding marked automata \mathcal{A}_M and
 856 \mathcal{B}_M , $\mathcal{L}(\mathcal{A}) \subseteq \mathcal{L}(\mathcal{B})$ if and only if $untime(\mathcal{L}(\mathcal{A}_M)) \subseteq untime(\mathcal{L}(\mathcal{B}_M))$.*

857 **Proof.** (\Rightarrow) Assume, $\mathcal{L}(\mathcal{A}) \subseteq \mathcal{L}(\mathcal{B})$ and assume there exists a word $w \in untime(\mathcal{L}(\mathcal{A}_M))$, but
 858 $w \notin untime(\mathcal{L}(\mathcal{B}_M))$. Now, there exists a timed word $\rho \in \mathcal{L}(\mathcal{A})$ such that, $f(\rho) = w$. Clearly,

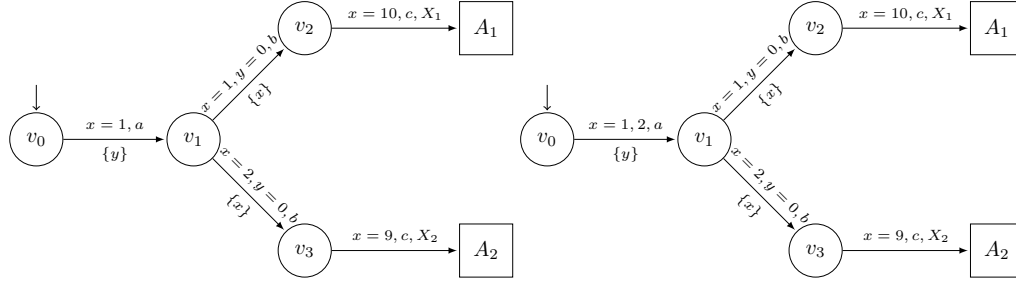


Figure 8 The automaton \mathcal{B} (left) and the faulty automaton $\mathcal{B}_{\mathcal{P}}$ (right)

859 $\rho \in \mathcal{L}(\mathcal{B})$, then clearly $f(\rho) = w \in \text{untimed}(\mathcal{L}(\mathcal{B}_M))$ a contradiction. So, $\text{untimed}(\mathcal{L}(\mathcal{A}_m)) \subseteq$
 860 $\text{untimed}(\mathcal{L}(\mathcal{B}_m))$.

861 (\Leftarrow) Assume, $\text{untimed}(\mathcal{L}(\mathcal{A}_M)) \subseteq \text{untimed}(\mathcal{L}(\mathcal{B}_M))$, and $\mathcal{L}(\mathcal{A}) \not\subseteq \mathcal{L}(\mathcal{B})$. Then, there
 862 exists a timed word $\rho \in \mathcal{L}(\mathcal{A})$ such that $\rho \notin \mathcal{L}(\mathcal{B})$. Assume $f(\rho) = w$, then clearly,
 863 $w \in \text{untimed}(\mathcal{L}(\mathcal{A}_M))$ and $w \in \text{untimed}(\mathcal{L}(\mathcal{B}_M))$. So, there exists a timed word $\rho' \in \mathcal{L}(\mathcal{A})$
 864 such that, $f(\rho') = w = f(\rho)$. According to Theorem 27 we can conclude that, $\rho \in \mathcal{L}(\mathcal{B})$ a
 865 contradiction. \blacktriangleleft

866 \blacktriangleright Remark 30. Lemma 29 shows that the timed and untimed language inclusion problems
 867 for IRTA are in fact the same problem. So, as we can solve the timed language inclusion
 868 problem by solving an untimed language inclusion problem of IRTA and vice-versa, the
 869 untimed language inclusion for IRTA is also EXPSPACE-Complete.

870 \blacktriangleright Theorem 31. Timed K - \forall -resilience of IRTA is EXPSPACE-Hard.

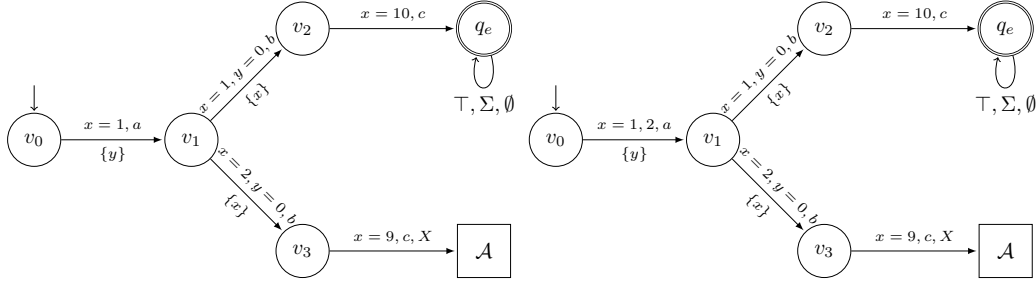
871 **Proof.** The proof is obtained by a reduction from the language inclusion problem of IRTA,
 872 known to be EXPSPACE-Complete [4]. The idea of the proof follows the same lines as the
 873 untimed K - \forall -resilience of timed automata. Assume we are given IRTA $\mathcal{A}_1, \mathcal{A}_2$. a, b, c are
 874 symbols not in the alphabets of $\mathcal{A}_1, \mathcal{A}_2$. Consider \mathcal{B} in Figure 8 (left). It is easy to see that
 875 $L(\mathcal{B}) = (a, 1)(b, 1)(c, 11)(L(\mathcal{A}_1) + 11)$, where $L(\mathcal{A}_1) + k = \{(a_1, d_1 + k)(a_2, d_2 + k) \dots (a_n, d_n + k) \mid$
 876 $(a_1, d_1) \dots (a_n, d_n) \in L(\mathcal{A}_1)\}$. Associate a fault model $\mathcal{P}(a) = 1$, where the fault of a is 1.
 877 We construct an IRTA $\mathcal{B}_{\mathcal{P}}$ as shown in Figure 8 (right). Notice that in general, IRTAs are
 878 not closed under the fault operation; the enlarged guard in \mathcal{B} would read $1 \leq x \leq 2$, and
 879 reset y . This transition violates the integer reset condition; however, since the transition
 880 on $1 < x < 2$ resetting y clearly does not lead to acceptance in $\mathcal{B}_{\mathcal{P}}$, we prune away that
 881 transition resulting in $\mathcal{B}_{\mathcal{P}}$ as in Figure 8 (right). Indeed, this resulting faulty automaton is
 882 an IRTA.

883 The language accepted by $\mathcal{B}_{\mathcal{P}}$ is $L(\mathcal{B}) \cup (a, 2)(b, 2)(c, 11)(L(\mathcal{A}_2) + 11)$. Considering $K = 2$,
 884 $\mathcal{B}_{\mathcal{P}}$ is BTN in 2 steps after the fault if and only if $L(\mathcal{A}_2) \subseteq L(\mathcal{A}_1)$. The EXPSPACE
 885 hardness of the timed K - \forall -resilience of IRTA follows from the EXPSPACE completeness of
 886 the inclusion of IRTA. \blacktriangleleft

887 \blacktriangleright Theorem 32. K - \exists -resilience for IRTA is PSPACE-Hard.

888 **Proof.** Consider an IRTA \mathcal{A} with alphabet Σ and the construction of an automata that
 889 uses a gadget shown below in Figure 9 (left). Let us call this automaton $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$. It
 890 is easy to see that the $L(\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}) = (a, 1)(b, 1)(c, 11)((\Sigma \times \mathbb{R})^* + 11)$, where $L(\mathcal{A}_1) + k =$
 891 $\{(a_1, d_1 + k)(a_2, d_2 + k) \dots (a_n, d_n + k) \mid (a_1, d_1) \dots (a_n, d_n) \in L(\mathcal{A}_1)\}$. The Σ loop on a
 892 particular accepting state q_e is responsible for acceptance of all timed word. Now, associate a

893 fault model $\mathcal{P}(a) \rightarrow 1$ with \mathcal{B} , where the fault of a is 1. Let us call this enlarged automaton
 894 $\mathcal{B}_{(\Sigma^* \subseteq \mathcal{A})_P}$. We can prune away the transition $1 < x < 2$ resetting y which does not lead
 895 to acceptance, and resulting in an IRTA with the same language, represented in Figure 9
 896 (right). The language accepted by $\mathcal{B}_{(\Sigma^* \subseteq \mathcal{A})_P}$ is $L(\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}) \cup (a, 2)(b, 2)(c, 11)(L(\mathcal{A}) + 11)$.
 897 The accepting states are $q_e \cup F$, where F is the set of final states of \mathcal{A} . Then $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ is
 898 K - \exists -resilient if and only if $L(\mathcal{A}) \neq \emptyset$. ◀



■ **Figure 9** The IRTA $\mathcal{B}_{\Sigma^* \subseteq \mathcal{A}}$ (left) and the faulty IRTA $\mathcal{B}_{(\Sigma^* \subseteq \mathcal{A})_P}$ (right)

899 ▶ **Remark 33.** The untimed language inclusion problem is shown to be EXPSPACE-Complete
 900 in Remark 30. The emptiness checking of timed automata is done by checking the emptiness
 901 of its untimed region automaton. So, to show the hardness of untimed K - \forall -resilient or
 902 K - \exists -resilient problems for IRTA, it is sufficient to reduce the untimed language inclusion
 903 problem and untimed language emptiness problem of IRTA respectively. This reduction can
 904 be done by using the same gadget as shown in Theorem 31 and Theorem 32 respectively.