Metric Temporal Description Logics with Interval-Rigid Names

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Abstract. In contrast to qualitative linear temporal logics, which can be used to state that some property will eventually be satisfied, metric temporal logics allow to formulate constraints on how long it may take until the property is satisfied. While most of the work on combining Description Logics (DLs) with temporal logics has concentrated on qualitative temporal logics, there has recently been a growing interest in extending this work to the quantitative case. In this paper, we complement existing results on the combination of DLs with metric temporal logics over the natural numbers by introducing interval-rigid names. This allows to state that elements in the extension of certain names stay in this extension for at least some specified amount of time.

1 Introduction

Description Logics [8] are a well-investigated family of logic-based knowledge representation languages, which provide the formal basis for the Web Ontology Language OWL.[1] As a consequence, DL-based ontologies are employed in many application areas, but they are particularly successful in the medical domain (see, e.g., the medical ontologies Galen and SNOMED CT[2]). For example, the concept of a patient with a concussion can formally be expressed in DLs as $\mathsf{Patient} \sqcap \exists \mathsf{finding}.\mathsf{Concussion}$, which is built from the concept names (i.e., unary predicates) $\mathsf{Patient}$ and $\mathsf{Concussion}$ and the role name (i.e., binary predicate) $\mathsf{finding}$ using the concept constructors conjunction ($\sqcap$) and existential restriction ($\exists r.C$). Concepts and roles can then be used within terminological and assertional axioms to state facts about the application domain, such as that concussion is a disease ($\mathsf{Concussion} \sqsubseteq \mathsf{Disease}$) and that patient Bob has a concussion ($\mathsf{Patient}(\mathsf{BOB}), \mathsf{finding}(\mathsf{BOB}, \mathsf{F1}), \mathsf{Concussion}(\mathsf{F1})$).

This example, taken from [9], can also be used to illustrate a shortcoming of pure DLs. For a doctor, it is important to know whether the concussed

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[1] https://www.w3.org/TR/2009/WD-owl2-overview-20090327/.
[2] See http://www.opengalen.org/ and http://www.snomed.org/.

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patient has lost consciousness, which is the reason why SNOMED CT contains a concept for "concussion with no loss of consciousness" [19]. However, the temporal pattern inherent in this concept (after the concussion, the patient remained conscious until the examination) cannot be modeled in the DL used for SNOMED CT.

To overcome the problem that pure DLs are not able to express such temporal patterns, a great variety of temporal extensions of DLs have been investigated in the literature.[3] In the present paper, we concentrate on the DL $\mathcal{ALC}$ and combine it with linear temporal logic (LTL), a point-based temporal logic whose semantics assumes a linear flow of time. But even if these two logics are fixed, there are several other design decisions to be made. One can either apply temporal operators only to axioms [9] or also use them within concepts [15,20]. With the latter, one can then formalize "concussion with no loss of consciousness" by the (temporal) concept

$$\exists\mathsf{finding}.\mathsf{Concussion} \sqcap (\mathsf{Conscious}\ \mathcal{U}\ \exists\mathsf{procedure}.\mathsf{Examination}),$$

where $\mathcal{U}$ is the *until*-operator of LTL. With the logic of [9], one cannot formulate temporal concepts, but could express that a particular patient, e.g., Bob, had a concussion and did not lose consciousness until he was examined. Another decision to be made is whether to allow for *rigid concepts and roles*, whose interpretation does not vary over time. For example, concepts like Human and roles like $\mathsf{hasFather}$ are clearly rigid, whereas $\mathsf{Conscious}$ and $\mathsf{finding}$ are flexible, i.e., not rigid. If temporal operators can be used within concepts, rigid concepts can be expressed using terminological axioms, but rigid roles cannot. In fact, they usually render the combined logic undecidable [15, Proposition 3.34]. In contrast, in the setting considered in [9], rigid roles do not cause undecidability, but adding rigidity leads to an increase in complexity.

In this paper, we address a shortcoming of the purely qualitative temporal description logics mentioned until now. The qualitative until-operator in our example does not say anything about how long after the concussion that examination happened. However, the above definition of "concussion with no loss of consciousness" is only sensible in case the examination took place in temporal proximity to the concussion. Otherwise, an intermediate loss of consciousness could also have been due to other causes. As another example, when formulating eligibility criteria for clinical trials, one needs to express quantitative temporal patterns [12] like the following: patients that had a treatment causing a reaction between 45 and 180 days after the treatment, and had no additional treatment before the reaction:

$$\mathsf{Treatment} \sqcap \bigcirc\big((\neg\mathsf{Treatment})\,\mathcal{U}_{[45,180]}\,\mathsf{Reaction}\big),$$

where $\bigcirc$ is the *next*-operator. On the temporal logic side, extensions of LTL by such intervals have been investigated in detail [1,2,16]. Using the next-operator of LTL as well as disjunction, their effect can actually be simulated within qualitative LTL, but if the interval boundaries are encoded in binary, this leads to

[3] We refer the reader to [15,17] for an overview of the field of temporal DLs.

an exponential blowup. The complexity results in [1] imply that this blowup can in general not be avoided, but in [16] it is shown that using intervals of a restricted form (where the lower bound is 0) does not increase the complexity compared to the qualitative case. In [13], the combination of the DL $\mathcal{ALC}$ with a metric extension of LTL is investigated. The paper considers both the case where temporal operators are applied only within concepts and the case where they are applied both within concepts and outside of terminological axioms. In Sect. 2, we basically recall some of the results obtained in [13], but show that they also hold if additionally temporalized assertional axioms are available.

In Sect. 3, we extend the logic $\text{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ of Sect. 2 with *interval-rigid* names, a means of expressiveness that has not been considered before. Basically, this allows one to state that elements belonging to a concept need to belong to that concept for at least k consecutive time points, and similarly for roles. For example, according to the WHO, patients with paucibacillary leprosy should receive MDT as treatment for 6 consecutive months,[4] which can be expressed by making the role $\mathsf{getMDTagainstPB}$ rigid for 6 time points (assuming that each time point represents one month). In Sect. 4, we consider the effect of adding interval-rigid concepts and roles as well as metric temporal operators to the logic $\mathcal{ALC}$-LTL of [9], where temporal operators can only be applied to axioms. Interestingly, in the presence of rigid roles, interval-rigid concepts actually cause undecidability. Without rigid roles, the addition of interval-rigid concepts and roles leaves the logic decidable, but in some cases increases the complexity (see Table 2). Finally, in Sect. 5 we investigate the complexity of this logic without interval-rigid names, which extends the analysis from [9] to quantitative temporal operators (see Table 3). An overview of the logics considered and their relations is shown in Fig. 1. Detailed proofs of all results can be found in [7].

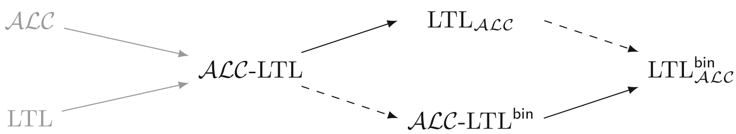

Fig. 1. Language inclusions, with languages investigated in this paper highlighted. Dashed arrows indicate same expressivity.

Related Work. Apart from the above references, we want to point out work on combining DLs with Halpern and Shoham's interval logic [3,4]. This setting is quite different from ours, since it uses intervals (rather than time points) as the basic time units. In [6], the authors combine $\mathcal{ALC}$ concepts with the (qualitative) operators $\Diamond$ ('at some time point') and $\Box$ ('at all time points') on roles, but do not consider quantitative variants. Recently, an interesting metric temporal extension of Datalog over the reals was proposed, which however cannot express interval-rigid names nor existential restrictions [11].

[4] See http://www.who.int/lep/mdt/duration/en/.

2 The Temporal Description Logic $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$

We first introduce the description logic $\mathcal{ALC}$ and its metric temporal extension $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ [13], which augments $\mathcal{ALC}$ by allowing metric temporal logic operators [1] both within $\mathcal{ALC}$ axioms and to combine these axioms. We actually consider a slight extension of $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ by assertional axioms, and show that this does not change the complexity of reasoning compared to the results of [13].

Syntax. Let $\mathsf{N_C}$, $\mathsf{N_R}$ and $\mathsf{N_I}$ be countably infinite sets of *concept names*, *role names*, and *individual names*, respectively. An $\mathcal{ALC}$ *concept* is an expression given by

$$C, D ::= A \mid \top \mid \neg C \mid C \sqcap D \mid \exists r.C,$$

where $A \in \mathsf{N_C}$ and $r \in \mathsf{N_R}$. $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *concepts* extend $\mathcal{ALC}$ concepts with the constructors $\bigcirc C$ and $C\,\mathcal{U}_I D$, where I is an interval of the form $[c_1, c_2]$ or $[c_1, \infty)$ with $c_1, c_2 \in \mathbb{N}$, $c_1 \leq c_2$, given in *binary*. We may use $[c_1, c_2)$ to abbreviate $[c_1, c_2 - 1]$, and similarly for the left endpoint. For example, $A\,\mathcal{U}_{[2,5)} B \sqcap \exists r.\bigcirc A$ is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ concept.

An $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *axiom* is either a *general concept inclusion (GCI)* of the form $C \sqsubseteq D$, or an *assertion* of the form $C(a)$ or $r(a, b)$, where C, D are $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ concepts, $r \in \mathsf{N_R}$, and $a, b \in \mathsf{N_I}$. $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *formulae* are expressions of the form

$$\phi, \psi ::= \alpha \mid \top \mid \neg\phi \mid \phi \wedge \psi \mid \bigcirc\phi \mid \phi\,\mathcal{U}_I \psi,$$

where α is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ axiom.

Semantics. A DL *interpretation* $\mathcal{I} = (\Delta^{\mathcal{I}}, \cdot^{\mathcal{I}})$ over a non-empty set $\Delta^{\mathcal{I}}$, called the *domain*, defines an *interpretation function* $\cdot^{\mathcal{I}}$ that maps each concept name $A \in \mathsf{N_C}$ to a subset $A^{\mathcal{I}}$ of $\Delta^{\mathcal{I}}$, each role name $r \in \mathsf{N_R}$ to a binary relation $r^{\mathcal{I}}$ on $\Delta^{\mathcal{I}}$ and each individual name $a \in \mathsf{N_I}$ to an element $a^{\mathcal{I}}$ of $\Delta^{\mathcal{I}}$, such that $a^{\mathcal{I}_i} \neq b^{\mathcal{I}_i}$ whenever $a \neq b$, $a, b \in \mathsf{N_I}$ (*unique name assumption*). As usual, we extend the mapping $\cdot^{\mathcal{I}}$ from concept names to $\mathcal{ALC}$ concepts as follows:

$$\top^{\mathcal{I}_i} := \Delta^{\mathfrak{I}}, \qquad (\neg C)^{\mathcal{I}_i} := \Delta^{\mathfrak{I}} \setminus C^{\mathcal{I}_i}, \qquad (C \sqcap D)^{\mathcal{I}_i} := C^{\mathcal{I}_i} \cap D^{\mathcal{I}_i},$$
$$(\exists r.C)^{\mathcal{I}_i} := \{d \in \Delta^{\mathfrak{I}} \mid \exists e \in C^{\mathcal{I}_i} : (d, e) \in r^{\mathcal{I}_i}\}.$$

A *(temporal DL) interpretation* is a structure $\mathfrak{I} = (\Delta^{\mathfrak{I}}, (\mathcal{I}_i)_{i \in \mathbb{N}})$, where each $\mathcal{I}_i = (\Delta^{\mathfrak{I}}, \cdot^{\mathcal{I}_i})$, $i \in \mathbb{N}$, is a DL interpretation over $\Delta^{\mathfrak{I}}$ (*constant domain assumption*) and $a^{\mathcal{I}_i} = a^{\mathcal{I}_j}$ for all $a \in \mathsf{N_I}$ and $i, j \in \mathbb{N}$, i.e., the interpretation of individual names is fixed. The mappings $\cdot^{\mathcal{I}_i}$ are extended to $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ concepts as follows:

$$(\bigcirc C)^{\mathcal{I}_i} := \{d \in \Delta^{\mathfrak{I}} \mid d \in C^{\mathcal{I}_{i+1}}\},$$
$$(C\,\mathcal{U}_I D)^{\mathcal{I}_i} := \{d \in \Delta^{\mathfrak{I}} \mid \exists k\colon k - i \in I,\ d \in D^{\mathcal{I}_k},\ \text{and } \forall j \in [i, k)\colon d \in C^{\mathcal{I}_j}\}.$$

The concept $C\,\mathcal{U}_I D$ requires D to be satisfied at some point in the interval I, and C to hold at all time points before that.

The *validity* of an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula ϕ in $\mathfrak{I}$ at time point $i \in \mathbb{N}$ (written $\mathfrak{I}, i \models \phi$) is inductively defined as follows:

$\mathfrak{I}, i \models C \sqsubseteq D$	iff $C^{\mathcal{I}_i} \subseteq D^{\mathcal{I}_i}$	$\mathfrak{I}, i \models \phi \wedge \psi$	iff $\mathfrak{I}, i \models \phi$ and $\mathfrak{I}, i \models \psi$
$\mathfrak{I}, i \models C(a)$	iff $a^{\mathcal{I}_i} \in C^{\mathcal{I}_i}$	$\mathfrak{I}, i \models \bigcirc\phi$	iff $\mathfrak{I}, i+1 \models \phi$
$\mathfrak{I}, i \models r(a,b)$	iff $(a^{\mathcal{I}_i}, b^{\mathcal{I}_i}) \in r^{\mathcal{I}_i}$	$\mathfrak{I}, i \models \phi\,\mathcal{U}_I\,\psi$	iff $\exists k\colon k - i \in I,\ \mathfrak{I}, k \models \psi$,
$\mathfrak{I}, i \models \neg\phi$	iff not $\mathfrak{I}, i \models \phi$		and $\forall j \in [i,k)\colon \mathfrak{I}, j \models \phi$.

As usual, we define $\bot := \neg\top$, $C \sqcup D := \neg(\neg C \sqcap \neg D)$, $\forall r.C := \neg(\exists r.\neg C)$, $\phi \vee \psi := \neg(\neg\phi \wedge \neg\psi)$, $\alpha\mathcal{U}\beta := \alpha\mathcal{U}_{[0,\infty)}\beta$, $\Diamond_I\alpha := \top\mathcal{U}_I\alpha$, $\Box_I\alpha := \neg\Diamond_I\neg\alpha$, $\Diamond\alpha := \top\mathcal{U}\alpha$, and $\Box\alpha := \neg\Diamond\neg\alpha$, where α, β are either concepts or formulae [8,15]. Note that, given the semantics of $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$, $\bigcirc\alpha$ is equivalent to $\Diamond_{[1,1]}\alpha$.

Relation to $\mathrm{LTL}_{\mathcal{ALC}}$. The notation $\cdot^{\mathsf{bin}}$ refers to the fact that the endpoints of the intervals are given in binary. However, this does not increase the expressivity compared to $\mathrm{LTL}_{\mathcal{ALC}}$ [17], where only the qualitative $\mathcal{U}$ operator is allowed. In fact, one can expand any formula $\phi\mathcal{U}_{[c_1,c_2]}\psi$ to $\bigvee_{c_1 \le i \le c_2}(\bigcirc^i\psi \wedge \bigwedge_{0 \le j < i}\bigcirc^j\phi)$, where $\bigcirc^i$ denotes i nested $\bigcirc$ operators, and similarly for concepts. Likewise, $\phi\mathcal{U}_{[c_1,\infty)}\psi$ is equivalent to $\left(\bigwedge_{0 \le i < c_1}\bigcirc^i\phi\right) \wedge \bigcirc^{c_1}\phi\mathcal{U}\psi$. If this transformation is recursively applied to subformulae, then the size of the resulting formula is exponential: ignoring the nested $\bigcirc$ operators, its syntax tree has polynomial depth and an exponential branching factor; and the $\bigcirc^i$ formulae have exponential depth, but introduce no branching. This blowup cannot be avoided in general [1,13].

Reasoning. We are interested in the complexity of the *satisfiability* problem in $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$, i.e., deciding whether there exists an interpretation $\mathfrak{I}$ such that $\mathfrak{I}, 0 \models \phi$ holds for a given $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula ϕ. We also consider a syntactic restriction from [9]: we say that ϕ is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *formula with global GCIs* if it is of the form $\Box\mathcal{T} \wedge \varphi$, where $\mathcal{T}$ is a conjunction of GCIs and φ is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula that does not contain GCIs. By *satisfiability w.r.t. global GCIs* we refer to the satisfiability problem restricted to such formulae.

First Results. The papers [13,17] consider the reasoning problems of concept satisfiability in $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ w.r.t. *TBoxes* (corresponding to formulae with global GCIs and without assertions) and satisfiability of $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *temporal TBoxes* (formulae without assertions). However, these results from [13,17] can be extended to our setting by incorporating *named types* into their quasimodel construction to deal with assertions (see also [20], our Sect. 3, and [15, Theorem 2.27]).

Theorem 1. *Satisfiability in* $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *is* 2-ExpSpace-*complete, and* ExpSpace-*complete w.r.t. global GCIs. In* $\mathrm{LTL}_{\mathcal{ALC}}$*, this problem is* ExpSpace-*complete, and* ExpTime-*complete w.r.t. global GCIs.*

Note that ExpSpace-completeness for $\mathrm{LTL}_{\mathcal{ALC}}$ with assertions has already been shown in [20]; we only state it here for completeness. In [13], also the intermediate logic $\mathrm{LTL}^{0,\infty}_{\mathcal{ALC}}$ was investigated, where only intervals of the form $[0, c]$ and $[c, \infty)$ are allowed. However, in [16], it was shown for a branching temporal logic that $\mathcal{U}_{[0,c]}$ can be simulated by the classical $\mathcal{U}$ operator, while only increasing the size of the formula by a polynomial factor. We extend this result to intervals of the form $[c, \infty)$, and apply it to $\mathrm{LTL}^{0,\infty}_{\mathcal{ALC}}$.

Theorem 2. *Any* $\mathrm{LTL}^{0,\infty}_{\mathcal{ALC}}$ *formula can be translated in polynomial time into an equisatisfiable* $\mathrm{LTL}_{\mathcal{ALC}}$ *formula.*

This reduction is quite modular; for example, if the formula has only global GCIs, then this is still the case after the reduction. In fact, the reduction applies to all sublogics of $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ that we consider in this paper. Hence, in the following we do not explicitly consider logics with the superscript $\cdot^{0,\infty}$, knowing that they have the same complexity as the corresponding temporal DLs using only $\mathcal{U}$.

3 $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ with Interval-Rigid Names

In many temporal DLs, so-called *rigid* names are considered, whose interpretation is not allowed to change over time. To formally define this notion, we fix a finite set $\mathsf{N_{Rig}} \subseteq \mathsf{N_C} \cup \mathsf{N_R}$ of *rigid* concept and role names, and require interpretations $\mathfrak{I} = (\Delta^{\mathfrak{I}}, (\mathcal{I}_i)_{i\in\mathbb{N}})$ to *respect* these names, in the sense that $X^{\mathcal{I}_i} = X^{\mathcal{I}_j}$ should hold for all $X \in \mathsf{N_{Rig}}$ and $i, j \in \mathbb{N}$. It turns out that $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ can already express rigid concepts via the (global) GCIs $C \sqsubseteq \bigcirc C$ and $\neg C \sqsubseteq \bigcirc\neg C$. The same does not hold for rigid roles, which lead to undecidability even in $\mathrm{LTL}_{\mathcal{ALC}}$ [15, Theorem 11.1]. Hence, it is not fruitful to consider rigid names in $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ (they will become meaningful later, when we look at other logics).

To augment the expressivity of temporal DLs while avoiding undecidability, we propose *interval-rigid* names. In contrast to rigid names, interval-rigid names only need to remain rigid for a limited period of time. Formally, we take a finite set $\mathsf{N_{IRig}} \subseteq (\mathsf{N_C} \cup \mathsf{N_R}) \setminus \mathsf{N_{Rig}}$ of *interval-rigid names*, and a function $\mathsf{iRig}\colon \mathsf{N_{IRig}} \to \mathbb{N}_{\geq 2}$. An interpretation $\mathfrak{I} = (\Delta^{\mathfrak{I}}, (\mathcal{I}_i)_{i\in\mathbb{N}})$ *respects* the interval-rigid names if the following holds for all $X \in \mathsf{N_{IRig}}$ with $\mathsf{iRig}(X) = k$, and $i \in \mathbb{N}$:

> For each $d \in X^{\mathcal{I}_i}$, there is a time point $j \in \mathbb{N}$ such that $i \in [j, j+k)$ and $d \in X^{\mathcal{I}_\ell}$ for all $\ell \in [j, j+k)$.

Intuitively, any element (or pair of elements) in the interpretation of an interval-rigid name must be in that interpretation for at least k consecutive time points. We call such a name *k-rigid*. The names in $(\mathsf{N_C} \cup \mathsf{N_R}) \setminus (\mathsf{N_{Rig}} \cup \mathsf{N_{IRig}})$ are called *flexible*. For simplicity, we assume that iRig assigns 1 to all flexible names.

We investigate the complexity of *satisfiability w.r.t. (interval-)rigid names* (or *(interval-)rigid concepts* if $\mathsf{N_{IRig}} \subseteq \mathsf{N_C}$ / $\mathsf{N_{Rig}} \subseteq \mathsf{N_C}$), which is defined as before, but considers only interpretations that respect (interval-)rigid names. Note that (interval-)rigid roles can be used to simulate (interval-)rigid concepts via existential restrictions $\exists r.\top$ (e.g., see [9]). Therefore, it is not necessary to consider the case where only role names can be (interval-)rigid. The fact that $\mathsf{N_{Rig}}$ and $\mathsf{N_{IRig}}$ are finite is not a restriction, as formulae can only use finitely many names. We assume that the values of iRig are given in binary.

Table 1 summarizes our results for $\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$. Since interval-rigid concepts A can be simulated by conjuncts of the form

$$\big(A \sqsubseteq \Box_{[0,k)}A\big) \wedge \Box\big(\neg A \sqsubseteq \bigcirc(\neg A \sqcup \Box_{[0,k)}A)\big),$$

Theorem 1 directly yields the complexity results in the right column (again, for sublogics of $\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$ this is not always so easy). The GCI $A \sqsubseteq \Box_{[0,k)}A$ that applies only to the first time point does not affect the complexity results, even if we restrict all other GCIs to be global.

Table 1. Complexity of satisfiability in $\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$ w.r.t. interval-rigid names. For (*), we have 2-ExpTime-completeness for the temporal semantics based on $\mathbb{Z}$ (Theorem 5).

	$\mathsf{N_{IRig}} \subseteq \mathsf{N_C} \cup \mathsf{N_R}$	$\mathsf{N_{IRig}} \subseteq \mathsf{N_C}$
$\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$	2-ExpSpace $\leq$ [Theorem 4]	2-ExpSpace $\geq$ [13]
$\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$, global GCIs	2-ExpTime-hard (*)	ExpSpace $\geq$ [2], $\leq$ [Theorem 1]
$\text{LTL}_{\mathcal{ALC}}$	2-ExpTime-hard	ExpSpace $\geq$ [15], $\leq$ [20]
$\text{LTL}_{\mathcal{ALC}}$, global GCIs	2-ExpTime-hard [Theorem 7]	ExpTime $\geq$ [18], $\leq$ [Theorem 1]

The complexity of $\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$ with interval-rigid roles is harder to establish. We first show in Sect. 3.1 that the general upper bound of 2-ExpSpace still holds, by a novel quasimodel construction. For global GCIs, we show 2-ExpTime-hardness in Sect. 4, by an easy adaption of a reduction from [9]. We show 2-ExpTime-completeness if we modify the temporal semantics to be infinite in both directions, i.e., replace $\mathbb{N}$ by $\mathbb{Z}$ in the definition of interpretations (see Sect. 3.2). We leave the case for the semantics based on $\mathbb{N}$ as future work. To simplify the proofs of the upper bounds, we usually assume that $\mathsf{N_{IRig}} \subseteq \mathsf{N_R}$ since interval-rigid concepts can be simulated. Moreover, for this section we assume that $\mathsf{N_{Rig}}$ is empty, as rigid concepts do not affect the complexity of $\text{LTL}^{\text{bin}}_{\mathcal{ALC}}$, and rigid roles make satisfiability undecidable.

3.1 Satisfiability Is in 2-ExpSpace

For the 2-ExpSpace upper bound, we extend the notion of *quasimodels* from [13]. In [13], quasimodels are abstractions of interpretations in which each time point is represented by a *quasistate*, which contains *types*. Each type describes the interpretation for a single domain element, while a quasistate collects the information about all domain elements at a single time point. Central for the complexity results in [13] is that every satisfiable formula has a quasimodel of a certain regular form, which can be guessed and checked in double exponential space. To handle interval-rigid roles, we extend this approach so that each quasistate additionally provides information about the temporal evolution of domain elements over a window of fixed width, and show that under this extended notion, satisfiability is still captured by the existence of regular quasimodels.

We now formalize this intuition. Let φ be an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula. Denote by $\mathsf{csub}(\varphi)/\mathsf{fsub}(\varphi)/\mathsf{ind}(\varphi)/\mathsf{rol}(\varphi)$ the set of all concepts/formulae/individuals/roles occurring in φ, by $\mathsf{cl}^{\mathsf{c}}(\varphi)$ the closure of $\mathsf{csub}(\varphi) \cup \{C\mathcal{U}D \mid C\mathcal{U}_{[c,\infty)}D \in \mathsf{csub}(\varphi)\}$ under single negations, and likewise for $\mathsf{cl}^{\mathsf{f}}(\varphi)$ and $\mathsf{fsub}(\phi)$. A *concept type for* φ is any subset t of $\mathsf{cl}^{\mathsf{c}}(\varphi) \cup \mathsf{ind}(\varphi)$ such that

T1 $\neg C \in t$ iff $C \notin t$, for all $\neg C \in \mathsf{cl}^{\mathsf{c}}(\varphi)$;
T2 $C \sqcap D \in t$ iff $C, D \in t$, for all $C \sqcap D \in \mathsf{cl}^{\mathsf{c}}(\varphi)$; and
T3 t contains at most one individual name.

Similarly, we define *formula types* $t \subseteq \mathsf{cl}^{\mathsf{f}}(\varphi)$ by the following conditions:

T1' $\neg\alpha \in t$ iff $\alpha \notin t$, for all $\neg\alpha \in \mathsf{cl}^{\mathsf{f}}(\varphi)$; and
T2' $\alpha \wedge \beta \in t$ iff $\alpha, \beta \in t$, for all $\alpha \wedge \beta \in \mathsf{cl}^{\mathsf{f}}(\varphi)$.

Intuitively, a concept type describes one domain element at a single time point, while a formula type expresses constraints on all domain elements. If $a \in t \cap \mathsf{ind}(\varphi)$, then t describes an named element, and we call it a *named type*.

To put an upper bound on the time window we have to look at, we consider the largest number occurring in φ and iRig, and denote it by ℓ_φ. Then, a *(concept/formula) run segment* for φ is a sequence $\sigma = \sigma(0) \dots \sigma(\ell_\varphi)$ composed exclusively of concept or formula types, respectively, such that

R1 $\bigcirc\alpha \in \sigma(0)$ iff $\alpha \in \sigma(1)$, for all $\bigcirc\alpha \in \mathsf{cl}^*(\varphi)$;
R2 for all $a \in \mathsf{ind}(\varphi)$ an $n \in (0, \ell_\varphi]$, we have $a \in \sigma(0)$ iff $a \in \sigma(n)$;
R3 for all $\alpha\mathcal{U}_I\beta \in \mathsf{cl}^*(\varphi)$, we have $\alpha\mathcal{U}_I\beta \in \sigma(0)$ iff (a) there is $j \in I \cap [0, \ell_\varphi]$ such that $\beta \in \sigma(j)$ and $\alpha \in \sigma(i)$ for all $i \in [0, j)$, or (b) I is of the form $[c, \infty)$ and $\alpha, \alpha\mathcal{U}\beta \in \sigma(i)$ for all $i \in [0, \ell_\varphi]$,

where cl^* is either cl^{c} or cl^{f} (as appropriate), and **R2** does not apply to formula run segments. A concept run segment captures the evolution of a domain element over a sequence of $\ell_\varphi + 1$ time points, and a formula run segment describes general constraints on the interpretation over a sequence of $\ell_\varphi + 1$ time points.

The evolution over the complete time line is captured by *(concept/formula) runs* for φ, which are infinite sequences $r = r(0)r(1)\dots$ such that each subsequence of length $\ell_\varphi + 1$ is a (concept/formula) run segment, and additionally

R4 $\alpha\mathcal{U}_{[c,\infty)}\beta \in r(n)$ implies that there is $j \geq n + c$ such that $\beta \in r(j)$ and $\alpha \in r(i)$ for all $i \in [n, j)$.

A concept run (segment) is *named* if it contains only (equivalently, any) named types. We may write r_a (σ_a) to denote a run (segment) that contains an individual name a. For a run (segment) σ, we write $\sigma^{>i}$ for the subsequence of σ starting at $i+1$, $\sigma^{<i}$ for the one stopping at $i-1$, and $\sigma^{[i,j]}$ for $\sigma(i) \dots \sigma(j)$.

Since we cannot explicitly represent infinite runs, we use run segments to construct them step-by-step. For this, it is important that a set of concept runs (segments) can actually be composed into a coherent model. In particular, we have to take care of (interval-rigid) role connections between elements. A *role constraint* for φ is a tuple (σ, σ', s, k), where σ, σ' are concept run segments, $s \in \mathsf{rol}(\varphi)$, and $k \in [1, \mathsf{iRig}(s)]$, such that

C1 $\{\neg C \mid \neg\exists s.C \in \sigma(0)\} \subseteq \sigma'(0)$; and
C2 if σ' is named, then σ is also named.

We write $\sigma \overset{s}{\underset{k}{\,}} \sigma'$ as a shorthand for the role constraint (σ, σ', s, k). Intuitively, $\sigma \overset{s}{\underset{k}{\,}} \sigma'$ means that the domain elements described by $\sigma(0), \sigma'(0)$ are connected by the role s at the current time point, and also at the $k-1$ previous time points. In this case, we need to ensure that these elements stay connected for at least the following $\mathsf{iRig}(s) - k$ time points. Condition **C1** ensures that, if $\sigma(0)$ cannot have any s-successors that satisfy C, then $\sigma'(0)$ does not satisfy C.

We can now describe the behaviour of a whole interpretation and its elements at a single time point, together with some bounded information about the future (up to ℓ_φ time points). A *quasistate* for φ is a pair $Q = (\mathcal{R}_Q, \mathcal{C}_Q)$, where $\mathcal{R}_Q$ is a set of run segments and $\mathcal{C}_Q$ a set of role constraints over $\mathcal{R}_Q$ such that

Q1 $\mathcal{R}_Q$ contains exactly one formula run segment σ_Q;
Q2 $\mathcal{R}_Q$ contains exactly one named run segment σ_a for each $a \in \mathsf{ind}(\varphi)$;
Q3 for all $C \sqsubseteq D \in \mathsf{cl}^{\mathsf{f}}(\varphi)$, we have $C \sqsubseteq D \in \sigma_Q(0)$ iff $C \in \sigma(0)$ implies $D \in \sigma(0)$ for all concept run segments $\sigma \in \mathcal{R}_Q$;
Q4 for all $C(a) \in \mathsf{cl}^{\mathsf{f}}(\varphi)$, we have $C(a) \in \sigma_Q(0)$ iff $C \in \sigma_a(0)$;
Q5 for all $s(a,b) \in \mathsf{cl}^{\mathsf{f}}(\varphi)$, we have $s(a,b) \in \sigma_Q(0)$ iff $\sigma_a \overset{s}{\underset{k}{\,}} \sigma_b \in \mathcal{C}_Q$ for some $k \in [1, \mathsf{iRig}(s)]$; and
Q6 for all $\sigma \in \mathcal{R}_Q$ and $\exists s.D \in \sigma(0)$, there is $\sigma \overset{s}{\underset{k}{\,}} \sigma' \in \mathcal{C}_Q$ with $D \in \sigma'(0)$ and $k \in [1, \mathsf{iRig}(s)]$.

We next capture when quasistates can be connected coherently to an infinite sequence. A pair (Q, Q') of quasistates is *compatible* if there is a *compatibility relation* $\pi \subseteq \mathcal{R}_Q \times \mathcal{R}_{Q'}$ such that

C3 every run segment in $\mathcal{R}_Q$ and $\mathcal{R}_{Q'}$ occurs at least once in the domain and range of π, respectively;
C4 each pair $(\sigma, \sigma') \in \pi$ satisfies $\sigma^{>0} = \sigma'^{<\ell_\varphi}$;
C5 for all $(\sigma_1, \sigma_1') \in \pi$ and $\sigma_1 \overset{s}{\underset{k}{\,}} \sigma_2 \in Q$ with $k < \mathsf{iRig}(s)$, there is $\sigma_1' \overset{s}{\underset{k+1}{\,}} \sigma_2' \in Q'$ with $(\sigma_2, \sigma_2') \in \pi$; and
C6 for all $(\sigma_1, \sigma_1') \in \pi$ and $\sigma_1' \overset{s}{\underset{k+1}{\,}} \sigma_2' \in Q'$ with $k > 1$, there is $\sigma_1 \overset{s}{\underset{k}{\,}} \sigma_2 \in Q$ with $(\sigma_2, \sigma_2') \in \pi$.

Such a relation makes sure that we can combine run segments of consecutive quasistates such that the interval-rigid roles are respected. Note that the unique formula run segments must be matched to each other, and likewise for the named run segments. Moreover, the set of all compatibility relations for a pair of quasistates (Q, Q') is closed under union, which means that compatible quasistates always have a unique maximal compatibility relation (w.r.t. set inclusion).

To illustrate this, consider Fig. 2, showing a sequence of pairwise compatible quasistates, each containing two run segments. Here, $\ell_\varphi = \mathsf{iRig}(s) = 3$. The relations π_0, π_1, and π_2 satisfy Conditions **C3**–**C6**, which, together with **C1** and **C2**, ensure that a run going through the types t_1, t_2, t_3, and t_4 can be connected to another run via the role s for at least 3 consecutive time points.

Finally, a *quasimodel* for φ is a pair $(S, \mathfrak{R})$, where S is an infinite sequence of compatible quasistates $S(0)S(1)\ldots$ and $\mathfrak{R}$ is a non-empty set of runs, such that

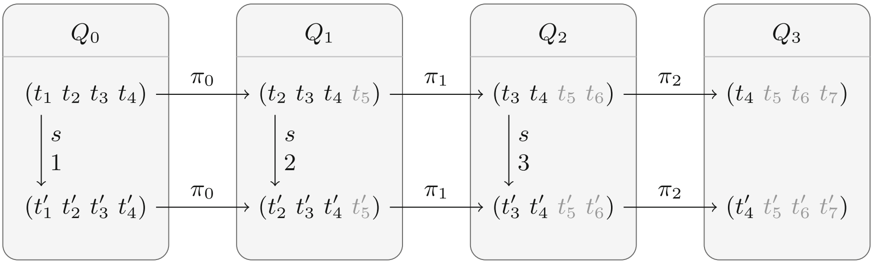

Fig. 2. Illustration of role constraints and compatibility relations.

M1 the runs in $\mathfrak{R}$ are of the form $\sigma_0(0)\sigma_1(0)\sigma_2(0)\dots$ such that, for every $i \in \mathbb{N}$, we have $(\sigma_i, \sigma_{i+1}) \in \pi_i$, where π_i is the maximal compatibility relation for the pair $(S(i), S(i+1))$;
M2 for every $\sigma \in \mathcal{R}_{S(i)}$, there exists a run $r \in \mathfrak{R}$ with $r^{[i,i+\ell_\varphi]} = \sigma$;
M3 every role constraint in $S(0)$ is of the form $\sigma_1 \overset{s}{\underset{1}{\downarrow}} \sigma_2$; and
M4 $\varphi \in \sigma_{S(0)}(0)$.

By **M1**, the runs $\sigma_0(0)\sigma_1(0)\sigma_2(0)\dots$ always contain the whole run segments $\sigma_0, \sigma_1, \sigma_2, \dots$, since we have $\sigma_1(0) = \sigma_0(1)$, $\sigma_2(0) = \sigma_0(2)$, and so on. Moreover, $\mathfrak{R}$ always contains exactly one formula run and one named run for each $a \in \mathsf{ind}(\varphi)$.

We can show that every quasimodel describes a satisfying interpretation for φ and, conversely, that every such interpretation can be abstracted to a quasimodel. Moreover, one can always find a quasimodel of a regular shape.

Lemma 3. *An* $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *formula* φ *is satisfiable w.r.t. interval-rigid names iff* φ *has a quasimodel* $(S, \mathfrak{R})$ *in which* S *is of the form*

$$S(0)\dots S(n)(S(n+1)\dots S(n+m))^\omega,$$

where n *and* m *are bounded triple exponentially in the size of* φ *and* iRig.

This allows us to devise a non-deterministic 2-ExpSpace algorithm that decides satisfiability of a given $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula. Namely, we first guess n and m, and then the quasistates $S(0), \dots, S(n+m)$ one after the other. To show that this sequence corresponds to a quasimodel as in Lemma 3, note that only three quasistates have to be kept in memory at any time, the sizes of which are double exponentially bounded in the size of the input: the current quasistate, the next quasistate, and the first repeating quasistate $S(n+1)$. 2-ExpSpace-hardness holds already for the case without interval-rigid names or assertions [13].

Theorem 4. *Satisfiability in* $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *with respect to interval-rigid names is* 2-ExpSpace*-complete.*

3.2 Global GCIs

For $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formulae with global GCIs, we can show a tight (2-ExpTime) complexity bound only if we consider a modified temporal semantics that uses $\mathbb{Z}$

instead of $\mathbb{N}$. With a semantics over $\mathbb{Z}$, every satisfiable formula has a quasimodel in which the unnamed run segments and role constraints are the same for all quasistates. This is not the case if the semantics is only defined for $\mathbb{N}$, since then a quasistate at time point 1 can have role constraints $\sigma \overset{s}{\underset{k}{\cdot}} \sigma'$ with $k > 1$, whereas one at time point 0 cannot (see **M3**).

Hence, interpretations are now of the form $\mathfrak{I} = (\Delta^{\mathfrak{I}}, (\mathcal{I}_i)_{i\in\mathbb{Z}})$, where $\Delta^{\mathfrak{I}}$ is a constant domain and $\mathcal{I}_i$ are classical DL interpretations, as before. Recall that an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula with global GCIs is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula of the form $\Box\mathcal{T} \wedge \phi$, where $\mathcal{T}$ is a conjunction of GCIs and ϕ is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula that does not contain GCIs. In order to enforce our GCIs on the whole time line (including the time points before 0), we replace $\Box\mathcal{T}$ with $\Box^-_+$ in that definition, where $\Box^-_+\mathcal{T}$ expresses that in all models $\mathfrak{I}$, $\mathfrak{I}, i \models \mathcal{T}$ for all $i \in \mathbb{Z}$. We furthermore slightly adapt some of the notions introduced in Sect. 3.1. First, to ensure that GCIs hold on the whole time line, we require (in addition to **T1'** and **T2'**) that all formula types contain all GCIs from $\mathcal{T}$. Additionally, we adapt the notions of runs $\ldots r(-1)r(0)r(1)\ldots$ and sequences $\ldots S(-1)S(0)S(1)\ldots$ of quasistates to be infinite in both directions. Hence, we can now drop Condition **M3**, reflecting the fact that, over $\mathbb{Z}$, role connections can exist before time point 0. All other definitions remain unchanged.

The complexity proof follows a similar idea as in the last section. We first show that every formula is satisfiable iff it has a quasimodel of a regular shape, which now is also constant in its unnamed part, in the sense that, if unnamed run segments and role constraints occur in $S(i)$, then they also occur in $S(j)$, for all $i, j \in \mathbb{Z}$. This allows us to devise an elimination procedure (in the spirit of [17, Theorem 3] and [13, Theorem 2]), with the difference that we eliminate run segments and role constraints instead of types, which gives us a 2-ExpTime upper bound. The matching lower bound can be shown similarly to Theorem 7 in Sect. 4.

Theorem 5. *Satisfiability in* $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ *w.r.t. interval-rigid names and global GCIs over* $\mathbb{Z}$ *is* 2-ExpTime*-complete.*

4 $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ with Interval-Rigid Names

After the very expressive DL $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$, we now focus on its sublogic $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$, which does not allow temporal operators within concepts (cf. [9]). That is, an $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ *formula* is an $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ formula in which all concepts are $\mathcal{ALC}$ concepts. Recall that $\mathcal{ALC}$-LTL, which has been investigated in [9] (though not with interval-rigid names), restricts $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ to intervals of the form $[0, \infty)$. In this section, we show several complexity lower bounds that already hold for $\mathcal{ALC}$-LTL with interval-rigid names. As done in [9], for brevity, we distinguish here the variants with global GCIs by the subscript $\cdot_{|gGCI}$. In contrast to $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$, in $\mathcal{ALC}$-LTL rigid concepts cannot be simulated by GCIs and rigid roles do not lead to undecidability [9]. Hence, we investigate here also the settings with rigid concepts and/or roles.

The results of this section are summarized in Table 2. Central to our hardness proofs is the insight that interval-rigid concepts can express the operator $\bigcirc$

on the concept level. In particular, we show that the combination of rigid roles with interval-rigid concepts already leads to undecidability, by a reduction from a tiling problem. If rigid names are disallowed, but we have interval-rigid names, we can only show 2-ExpTime-hardness. If only interval-rigid concepts are allowed, then satisfiability is ExpSpace-hard. All of these hardness results already hold for $\mathcal{ALC}$-LTL, and some of them even with global GCIs.

Table 2. Complexity of satisfiability in $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ w.r.t. (interval-)rigid names.

	$\mathsf{N_{IRig}} \subseteq \mathsf{N_C}$, $\mathsf{N_{Rig}} \subseteq \mathsf{N_C} \cup \mathsf{N_R}$	$\mathsf{N_{IRig}} \subseteq \mathsf{N_C} \cup \mathsf{N_R}$, $\mathsf{N_{Rig}} \subseteq \mathsf{N_C}$ or $\mathsf{N_{Rig}} = \emptyset$	$\mathsf{N_{IRig}} \subseteq \mathsf{N_C}$, $\mathsf{N_{Rig}} \subseteq \mathsf{N_C}$ or $\mathsf{N_{Rig}} = \emptyset$
$\mathcal{ALC}$-LTL$^{\mathsf{bin}}$	Undec.	2-ExpTime-hard	ExpSpace $\leq$ [Theorem 1]
$\mathcal{ALC}$-LTL$^{\mathsf{bin}}_{\vert gGCI}$	Undec.	2-ExpTime-hard	ExpSpace $=$ [2]
$\mathcal{ALC}$-LTL	Undec.	2-ExpTime-hard	ExpSpace $\geq$ [Theorem 8]
$\mathcal{ALC}$-LTL$_{\vert gGCI}$	Undec. [Theorem 6]	2-ExpTime-hard [Theorem 7]	ExpTime $\geq$ [18], $\leq$ [Theorem 1]

4.1 Rigid Roles and Interval-Rigid Concepts

We show that satisfiability of $\mathcal{ALC}$-LTL with rigid roles and interval-rigid concepts is undecidable, even if we only allow global GCIs. Our proof is by a reduction from the following tiling problem.

> Given a finite set of tile types T with horizontal and vertical compatibility relations H and V, respectively, and $t_0 \in T$, decide whether one can tile $\mathbb{N} \times \mathbb{N}$ with t_0 appearing infinitely often in the first row.

We define an $\mathcal{ALC}$-LTL$_{\vert gGCI}$ formula ϕ_T that expresses this property. In our encoding, we use the following names:

- a rigid role name r to encode the vertical dimension of the $\mathbb{N} \times \mathbb{N}$ grid;
- flexible concept names A^0, A^1, A^2 to encode the progression along the horizontal (temporal) dimension; for convenience, we consider all superscripts modulo 3, i.e., we have $A^3 = A^0$ and $A^{-1} = A^2$;
- flexible concept names P_t, $t \in T$, to denote the current tile type;
- 2-rigid concept names N_t^0, N_t^1, N_t^2, for the horizontally adjacent tile type;
- an individual name a denotes the first row of the grid.

We define ϕ_T as the conjunction of the following $\mathcal{ALC}$-LTL$_{\vert gGCI}$ formulae. First, every domain element must have exactly one tile type:

$$\Box\Big(\top \sqsubseteq \bigsqcup_{t\in T}\Big(P_t \sqcap \bigsqcap_{t'\in T,\, t\neq t'} \neg P_{t'}\Big)\Big)$$

For the vertical dimension, we enforce an infinite rigid r-chain starting from a, and restrict adjacent tile types to be compatible:

$$\Box(\top \sqsubseteq \exists r.\top), \quad \Box\Big(P_t \sqsubseteq \bigsqcup_{(t,t')\in V} \forall r.P_{t'}\Big)$$

For each time point i, we mark all individuals along the r-chain with the concept name $A^{(i \bmod 3)}$, by using the following formulae, for $0 \le i \le 2$:

$$A^0(a), \quad \Box\big(A^i(a) \to \bigcirc A^{i+1}(a)\big), \quad \Box(A^i \sqsubseteq \neg A^{i+1} \sqcap \forall r.A^i)$$

To encode the compatibility of horizontally adjacent tiles, we add the following formulae, for $0 \le i \le 2$ and $t \in T$:

$$\Box\Big(P_t \sqcap A^i \sqsubseteq \bigsqcup_{(t,t')\in H} N^i_{t'}\Big), \quad \Box(N^i_t \sqcap A^{i+1} \sqsubseteq P_t), \quad \Box(A^{i-1} \sqsubseteq \neg N^i_t)$$

These express that any domain element with tile type t (expressed by P_t) at a time point marked with A^i must have a compatible type t' at the next time point (expressed by $N^i_{t'}$). Since all $N^i_{t'}$ are false at the previous time point (designated by A^{i-1}) and $\mathsf{iRig}(N^i_{t'}) = 2$, any $N^i_{t'}$ that holds at the current time point is still active at the next time point (described by A^{i+1}), where it then implies $P_{t'}$.

Finally, we express the condition on t_0 via the formula $\Box\Diamond P_{t_0}(a)$. We now obtain the claimed undecidability from known results about the tiling problem [14].

Theorem 6. *Satisfiability in $\mathcal{ALC}$-LTL$_{|gGCI}$ w.r.t. rigid roles and interval-rigid concepts is Σ^1_1-hard, and thus not even recursively enumerable.*

4.2 Interval-Rigid Roles

Since rigid roles cause undecidability, we consider the case where instead only interval-rigid roles (and concepts) are allowed, and obtain 2-ExpTime-hardness by an easy adaptation of a result for $\mathcal{ALC}$-LTL$_{|gGCI}$ with rigid roles from [9].

Theorem 7. *Satisfiability in $\mathcal{ALC}$-LTL$_{|gGCI}$ with respect to interval-rigid names is* 2-ExpTime*-hard.*

4.3 Rigid and Interval-Rigid Concepts

As the last setting, we consider the case where only concept names can be rigid or interval-rigid, and show ExpSpace-completeness. For the upper bound, recall from Sect. 3 that rigid concepts and interval-rigid concepts are expressible in LTL$^{0,\infty}_{\mathcal{ALC}}$ via global GCIs, so that we can apply Theorem 1. The same observation yields an ExpTime upper bound for satisfiability in $\mathcal{ALC}$-LTL w.r.t. global GCIs, which is tight since satisfiability in ordinary $\mathcal{ALC}$ is already ExpTime-hard [18].

We show the ExpSpace lower bound by a reduction from satisfiability of $\mathcal{ALC}$-LTL$^{\bigcirc}$, the extension of $\mathcal{ALC}$-LTL in which $\bigcirc$ can be applied to concepts, to satisfiability of $\mathcal{ALC}$-LTL w.r.t. interval-rigid concepts. It is shown in [15, Theorem 11.33] that satisfiability in (a syntactic variant of) $\mathcal{ALC}$-LTL$^{\bigcirc}$ is ExpSpace-hard. To simulate $\bigcirc$ using interval-rigid concept names, we use a similar construction as in Sect. 4.1, where we mark all individuals at time point i with $A^{(i \bmod 3)}$, and use 2-rigid concept names to transfer information between time points. More precisely, we first define an $\mathcal{ALC}$-LTL formula ψ as the conjunction of the following formulae, where $0 \leq i \leq 2$:

$$(\top \sqsubseteq A^0), \quad \Box((\top \sqsubseteq A^i) \to \bigcirc(\top \sqsubseteq A^{i+1})), \quad \Box(A^i \sqsubseteq \neg A^{i+1})$$

We now simulate concepts of the form $\bigcirc C$ via fresh, 2-rigid concept names $A^i_{\bigcirc C}$, $0 \leq i \leq 2$. Given any $\mathcal{ALC}$-LTL$^{\bigcirc}$ formula α (resp., $\mathcal{ALC}$-LTL$^{\bigcirc}$ concept D), we denote by $\alpha^{\bigcirc}$ (resp., $D^{\bigcirc}$) the result of replacing each outermost concept of the form $\bigcirc C$ in α (resp., D) by

$$\bigsqcup_{0 \leq i \leq 2} (A^i_{\bigcirc C} \sqcap A^i).$$

To express the semantics of $\bigcirc C$, we use the conjunction $\psi_{\bigcirc C}$ of the following formulae (where the replacement operator $\cdot^{\bigcirc}$ is applied to the inner concept C):

$$\Box(A^i_{\bigcirc C} \sqcap A^{i+1} \sqsubseteq C^{\bigcirc}), \quad \Box(C^{\bigcirc} \sqcap A^{i+1} \sqsubseteq A^i_{\bigcirc C}), \quad \Box(A^{i-1} \sqsubseteq \neg A^i_{\bigcirc C})$$

As in Sect. 4.1, $A^i_{\bigcirc C}$ must either be satisfied at both time points designated by A^i and A^{i+1}, or at neither of them. Furthermore, an individual satisfies $\bigcirc C$ iff it satisfies $A^i_{\bigcirc C} \sqcap A^i$ for some i, $0 \leq i \leq 2$. One can show that an $\mathcal{ALC}$-LTL$^{\bigcirc}$ formula ϕ is satisfiable iff the $\mathcal{ALC}$-LTL formula $\phi^{\bigcirc} \wedge \psi \wedge \bigwedge_{\bigcirc C \in \mathsf{csub}(\phi)} \psi_{\bigcirc C}$ is satisfiable.

Theorem 8. *Satisfiability in $\mathcal{ALC}$-LTL with respect to interval-rigid concepts is* ExpSpace*-hard.*

5 $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ Without Interval-Rigid Names

To conclude our investigation of metric temporal DLs, we consider the setting of $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ without interval-rigid names. Table 3 summarizes the results of this section, where we also include the known results about $\mathcal{ALC}$-LTL for comparison [9]. Observe that all lower bounds follow from known results. In particular, ExpSpace-hardness for $\mathcal{ALC}$-LTL$^{\mathsf{bin}}_{|gGCI}$ is inherited from LTL$^{\mathsf{bin}}$ [1,2], while rigid role names increase the complexity to 2-ExpTime in $\mathcal{ALC}$-LTL$_{|gGCI}$ [9].

The upper bounds can be shown using a unified approach that was first proposed in [9]. The idea is to split the satisfiability test into two parts: one for the temporal and one for the DL dimension. In what follows, let ϕ be an $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ formula. The *propositional abstraction* ϕ^{p} is the propositional LTL$^{\mathsf{bin}}$ formula obtained from ϕ by replacing every $\mathcal{ALC}$ axiom by a propositional variable

Table 3. Complexity of satisfiability in $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ without interval-rigid names.

	$\mathsf{N_{Rig}} \subseteq \mathsf{N_C} \cup \mathsf{N_R}$	$\mathsf{N_{Rig}} \subseteq \mathsf{N_C}$	$\mathsf{N_{Rig}} = \emptyset$
$\mathcal{ALC}$-LTL$^{\mathsf{bin}}$	2-ExpTime $\leq$ [Theorem 10]	ExpSpace $\leq$ [Theorem 10]	ExpSpace
$\mathcal{ALC}$-LTL$^{\mathsf{bin}}_{\vert gGCI}$	2-ExpTime	ExpSpace	ExpSpace $\geq$ [1]
$\mathcal{ALC}$-LTL	2-ExpTime	NExpTime [9]	ExpTime $\leq$ [9]
$\mathcal{ALC}$-LTL$_{\vert gGCI}$	2-ExpTime $\geq$ [9]	ExpTime $\leq$ [9]	ExpTime $\geq$ [18]

in such a way that there is a 1:1 relationship between the $\mathcal{ALC}$ axioms $\alpha_1, \ldots, \alpha_m$ occurring in ϕ and the propositional variables $p_1, \ldots, p_m$ in ϕ^{p}.

The goal is to try to find a model of ϕ^{p} and then use it to construct a model of ϕ (if such a model exists). While satisfiability of ϕ implies that ϕ^{p} is also satisfiable, the converse is not true. For example, the propositional abstraction $p \wedge q \wedge \neg r$ of $\phi = A \sqsubseteq B \wedge A(a) \wedge \neg B(a)$ is satisfiable, while ϕ is not. To rule out such cases, we collect the propositional worlds occurring in a model of ϕ^{p} into a (non-empty) set $\mathcal{W} \subseteq 2^{\{p_1,\ldots,p_m\}}$, which is then used to check the satisfiability of the original formula (w.r.t. rigid names). This is captured by the LTL$^{\mathsf{bin}}$ formula $\phi^{\mathrm{p}}_{\mathcal{W}} := \phi^{\mathrm{p}} \wedge \phi_{\mathcal{W}}$, where $\phi_{\mathcal{W}}$ is the (exponential) LTL formula

$$\Box \bigvee_{W \in \mathcal{W}} \left(\bigwedge_{p \in W} p \wedge \bigwedge_{p \in \overline{W}} \neg p \right)$$

in which $\overline{W} := \{p_1, \ldots, p_m\} \setminus W$ denotes the complement of W. The formula $\phi^{\mathrm{p}}_{\mathcal{W}}$ states that, when looking for a propositional model of ϕ^{p}, we are only allowed to use worlds from $\mathcal{W}$.

Since satisfiability of ϕ implies satisfiability of $\phi^{\mathrm{p}}_{\mathcal{W}}$ for some $\mathcal{W}$, we can proceed as follows: choose a set of worlds $\mathcal{W}$, test whether $\phi^{\mathrm{p}}_{\mathcal{W}}$ is satisfiable, and then check whether a model with worlds from $\mathcal{W}$ can indeed be lifted to a temporal DL interpretation (respecting rigid names). To check the latter, we consider the conjunction $\bigwedge_{p_j \in W} \alpha_j \wedge \bigwedge_{p_j \in \overline{W}} \neg \alpha_j$ for every $W \in \mathcal{W}$. However, the rigid names require that all these conjunctions are *simultaneously* checked for satisfiability. To tell apart the *flexible* names X occurring in different elements of $\mathcal{W} = \{W_1, \ldots, W_k\}$, we introduce copies $X^{(i)}$ for all $i \in [1, k]$. The axioms $\alpha_j^{(i)}$ are obtained from α_j by replacing every flexible name X by $X^{(i)}$, which yields the following conjunction of exponential size:

$$\chi_{\mathcal{W}} := \bigwedge_{i=1}^{k} \Big(\bigwedge_{p_j \in W_i} \alpha_j^{(i)} \wedge \bigwedge_{p_j \in \overline{W_i}} \neg \alpha_j^{(i)} \Big).$$

The following characterization from [9] can be easily adapted to our setting:

Lemma 9 (Adaptation of [9]). *An $\mathcal{ALC}$-LTL$^{\mathsf{bin}}$ formula ϕ is satisfiable w.r.t. rigid names iff a set $\mathcal{W} \subseteq 2^{\{p_1,\ldots,p_m\}}$ exists so that $\phi^{\mathrm{p}}_{\mathcal{W}}$ and $\chi_{\mathcal{W}}$ are both satisfiable.*

To obtain the upper bounds in Table 3, recall from Sect. 2 that there is an exponentially larger LTL formula $\phi^{\mathrm{p}\prime}$ that is equivalent to the $\mathrm{LTL}^{\mathsf{bin}}$ formula ϕ^{p}. Since $\phi_{\mathcal{W}}$ is also an LTL formula of exponential size, satisfiability of the conjunction $\phi^{\mathrm{p}\prime} \wedge \phi_{\mathcal{W}}$ can be checked in ExpSpace. Since the complexity of the satisfiability problem for $\chi_{\mathcal{W}}$ remains the same as in the case of $\mathcal{ALC}$-LTL, we obtain the claimed upper bounds from the techniques in [9]. This means that, in most cases, the complexity of the DL part is dominated by the ExpSpace complexity of the temporal part. The only exception is the 2-ExpTime-bound for $\mathcal{ALC}$-$\mathrm{LTL}^{\mathsf{bin}}$ with rigid names.

Theorem 10. *Satisfiability in $\mathcal{ALC}$-$\mathrm{LTL}^{\mathsf{bin}}$ is in* 2-ExpTime *w.r.t. rigid names, and in* ExpSpace *w.r.t. rigid concepts.*

6 Conclusions

We investigated a series of extensions of $\mathrm{LTL}_{\mathcal{ALC}}$ and $\mathcal{ALC}$-LTL with interval-rigid names and metric temporal operators, with complexity results ranging from ExpTime to 2-ExpSpace. Some cases were left open, such as the precise complexity of $\mathrm{LTL}^{\mathsf{bin}}_{\mathcal{ALC}}$ with global GCIs, for which we have a partial result for the temporal semantics based on $\mathbb{Z}$. Nevertheless, this paper provides a comprehensive guide to the complexities faced by applications that want to combine ontological reasoning with quantitative temporal logics.

In principle, the arguments for $\mathcal{ALC}$-$\mathrm{LTL}^{\mathsf{bin}}$ in Sect. 5 are also applicable if we replace $\mathcal{ALC}$ by the light-weight DLs *DL-Lite* or $\mathcal{EL}$, yielding tight complexity bounds based on the known results from [5,10]. It would be interesting to investigate temporal DLs based on *DL-Lite* and $\mathcal{EL}$ with interval-rigid roles and metric operators.

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