Proceedings of the Twelfth International Conference on the Principles of Knowledge Representation and Reasoning (KR 2010)

On the Complexity of Axiom Pinpointing in the $\mathcal{EL}$ Family of Description Logics

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Abstract

We investigate the computational complexity of axiom pinpointing, which is the task of finding minimal subsets of a Description Logic knowledge base that have a given consequence. We consider the problems of enumerating such subsets with and without order, and show hardness results that already hold for the propositional Horn fragment, or for the Description Logic $\mathcal{EL}$. We show complexity results for several other related decision and enumeration problems for these fragments that extend to more expressive logics. In particular we show that hardness of these problems depends not only on expressivity of the fragment but also on the shape of the axioms used.

Introduction

Description Logics (DLs) (Baader et al. 2003) are a well-established family of logic-based knowledge representation formalisms that are used to represent the conceptual knowledge of an application domain in a structured and formally well-understood way. DLs have proven successful in various application domains, but they have gained increased attention due to the fact that they provide the logical underpinning of OWL (Horrocks, Patel-Schneider, and van Harmelen 2003), the standard ontology language for the semantic web. As a consequence of this standardization, several ontology editors (Knublauch et al. 2004; Kalyanpur et al. 2006; Horridge, Tsarkov, and Redmond 2006), now support OWL and ontologies, or knowledge bases (KBs), written in OWL are employed in more and more applications. As the sizes of these KBs grow, tools that support knowledge engineers in maintaining their quality become more important. In real world applications often the knowledge engineer not only wants to know whether her KB has a certain (unwanted) consequence or not, but also wants to know *why* it has this consequence. Even for KBs of moderate size, finding explanations for a given a consequence is not an easy task without getting support from an automated tool. The task of finding explanations for a given consequence, i.e., minimal subsets of the original KB that have the given consequence is called *axiom pinpointing* in the literature.

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Existing work on axiom pinpointing in DLs can be classified under two main categories, namely the glass-box approach, and the black-box approach. The idea underlying the *glass-box approach* is to extend the existing reasoning algorithms so that they can keep track of the axioms in the KB, and detect which of these axioms are responsible for a given consequence. In (Schlobach and Cornet 2003) a pinpointing extension of the well-known tableau-based satisfiability algorithm for the DL $\mathcal{ALC}$ (Schmidt-Schauß and Smolka 1991) has been introduced. Later in (Parsia, Sirin, and Kalyanpur 2005), this approach has been further extended to DLs that are more expressive than $\mathcal{ALC}$. In (Meyer et al. 2006) a pinpointing algorithm for $\mathcal{ALC}$ with general concept inclusions (GCIs) has been presented by following the approach in (Baader and Hollunder 1995). In order to overcome the problem of developing a pinpointing extension for every particular tableau-based algorithm, a general pinpointing extension for tableau algorithms has been developed in (Baader and Peñaloza 2007; 2010b). Similarly, an automata-based general approach for obtaining glass-box pinpointing algorithms has been introduced in (Baader and Peñaloza 2008; 2010a).

In contrast to the glass-box approach, the idea underlying the *black-box approach* is to make use of the existing highly optimized reasoning algorithms. The most naïve black-box approach would of course be to generate every subset of the originial KB, and ask a DL reasoner whether this subset has the given consequence or not, which obviously is very inefficient. In (Kalyanpur et al. 2007; Suntisrivaraporn et al. 2008) more efficient approaches based on Reiter's hitting set tree algorithm (Reiter 1987) have been presented. The experimental resuts in (Kalyanpur et al. 2007) demonstrate that this approach behaves quite well in practice on realistic KBs written in expressive DLs. A similar approach has successfully been used in (Horridge, Parsia, and Sattler 2009) for explaining inconsistencies in OWL ontologies. The main advantages of the black-box approach are that one can use existing DL reasoners, and that it is independent of the DL reasoner being used. In (Horridge, Parsia, and Sattler 2008) the black-box approach has been used for computing more fine grained explanations, i.e., not just the set of relevant axioms in the KB but parts of these axioms that actually lead to the given consequence.

Although various methods and aspects of axiom pinpointing have been considered in the literature, its computational complexity has not been investigated in detail yet. Obviously, axiom pinpointing is at least as hard as reasoning. Nevertheless, especially for tractable DLs it makes sense to investigate whether explanations for a consequence can efficiently be enumerated or not. In (Baader, Peñaloza, and Suntisrivaraporn 2007) it has been shown that a given consequence can have exponentially-many explanations (there called *MinAs*, which stands for *minimal axiom sets*), and checking the existence of a MinA within a cardinality bound is NP-complete. There it has also been shown that in a setting where MinAs are required to contain certain (static) part of the KB, then the set of all MinAs cannot be computed in output polynomial time. In (Peñaloza and Sertkaya 2009) among other results we have shown that without the static part this problem is at least as hard as computing minimal transversals of a hypergraph. We have also shown that if the MinAs are required to be output in a specified order, then the problem is not polynomial delay. In (Sebastiani and Vescovi 2009) a promising method that uses modern conflict-driven SAT solvers for axiom pinpointing in $\mathcal{EL}$ has been presented. The method roughly consists of generating propositional Horn formulas representing part or all the deduction steps performed by a classification algorithm, and manipulating them by the help of a SAT solver for computing a single MinA or for computing all MinAs.

In the present paper we present several new interesting complexity results on axiom pinpointing. We give a polynomial delay algoritm for enumerating MinAs in the Horn setting, show that for dual-Horn KBs the problem is at least as hard as hypergraph transversal enumeration, and for $\mathcal{EL}$ KBs it is not output polynomial. We show that if MinAs are required to be output in a specified order, then for dual-Horn and $\mathcal{EL}$ KBs this cannot be done with polynomial delay. We also consider several other decision and enumeration problems on MinAs in different settings.

Preliminaries

We briefly recall basic notions from propositional logic, DLs, and complexity of enumeration. In propositional logic we build formulae using a set of *propositional variables* and the *Boolan connectives* $\neg$ (negation), $\vee$ (disjunction) and $\wedge$ (conjunction). A variable or its negation is called a *literal*, and a disjunction of literals, e.g. $\neg p_1 \vee \neg p_2 \vee p_3$ is called a *clause*. Clauses can also be written as *implications* of the form $p_1 \wedge p_2 \rightarrow p_3$. A clause is called a *Horn (dual-Horn)* clause if it contains at most one positive (negative) literal, and a *definite Horn (dual-Horn)* clause if it contains exactly one positive (negative) literal. Throughout the text we will call definite Horn (dual-Horn) clauses just Horn (dual-Horn) clauses for short. We will call clauses with exactly one positive and one negative literal like $p_1 \rightarrow p_2$ as *core* clauses.

In DLs one formalizes the relevant notions of an application domain with *concept descriptions*. Concept descriptions are inductively built with the help of a set of *constructors*, starting with a set $\mathsf{N_C}$ of *concept names* and a set $\mathsf{N_R}$ of *role names*. $\mathcal{EL}$ concept descriptions are formed using the three constructors $\sqcap, \exists$ and $\top$ as shown in the upper part of Table 1. An $\mathcal{EL}$ *TBox* is a finite set of *general*

Syntax	Semantics
$\top$	$\Delta^{\mathcal{I}}$
$C \sqcap D$	$C^{\mathcal{I}} \cap D^{\mathcal{I}}$
$\exists r.C$	$\{x \in \Delta^{\mathcal{I}} \mid \exists y \in \Delta^{\mathcal{I}} : (x,y) \in r^{\mathcal{I}} \wedge y \in C^{\mathcal{I}}\}$
$C \sqsubseteq D$	$C^{\mathcal{I}} \subseteq D^{\mathcal{I}}$

Table 1: Syntax and semantics of $\mathcal{EL}$.

concept inclusion axioms, or *GCIs*, whose syntax is shown in the lower part of Table 1. The semantics of $\mathcal{EL}$ is defined in terms of *interpretations* $\mathcal{I} = (\Delta^{\mathcal{I}}, \cdot^{\mathcal{I}})$, where the *domain* $\Delta^{\mathcal{I}}$ is a non-empty set of individuals, and the interpretation function $\cdot^{\mathcal{I}}$ maps each concept name $A \in \mathsf{N_C}$ to a subset $A^{\mathcal{I}}$ of $\Delta^{\mathcal{I}}$ and each role name $r \in \mathsf{N_R}$ to a binary relation $r^{\mathcal{I}}$ on $\Delta^{\mathcal{I}}$. The mapping $\cdot^{\mathcal{I}}$ can be extended to arbitrary concept descriptions as shown in the second colum of Table 1. An interpretation $\mathcal{I}$ is a *model* of a TBox $\mathcal{T}$ if, for every GCI in $\mathcal{T}$ the conditions on the semantics column of Table 1 are satisfied. The main inference problem for $\mathcal{EL}$ is the subsumption problem (Baader 2003; Brandt 2004): given two $\mathcal{EL}$ concept descriptions C, D and an $\mathcal{EL}$ TBox $\mathcal{T}$, check if C is *subsumed* by D w.r.t. $\mathcal{T}$ (written $\mathcal{T} \models C \sqsubseteq D$), i.e, check if $C^{\mathcal{I}} \subseteq D^{\mathcal{I}}$ holds in every model $\mathcal{I}$ of $\mathcal{T}$. We will call a concept description *simple* if it is of the form A or $\exists r.A$ for $A \in \mathsf{N_C}, r \in \mathsf{N_R}$, and a GCI a *Horn-$\mathcal{EL}$ GCI* if it is of the form $C_1 \sqcap \ldots \sqcap C_n \sqsubseteq D$, where C_i, D are simple concept descriptions, $1 \leq i \leq n$.

We will refer to both propositional clauses and $\mathcal{EL}$ GCIs as *axioms*, and a set of axioms as a *knowledge base (KB)*. We will say that a KB is a Horn (core, dual-Horn, Horn-$\mathcal{EL}$, $\mathcal{EL}$) KB if it contains only Horn (core, dual-Horn, Horn-$\mathcal{EL}$, $\mathcal{EL}$) axioms. We are going to formulate our problems in a generic way without referring to a specific type of KB, and show our results for each KB type separately.

Note that core axioms are a special case of both, Horn and dual-Horn axioms. Likewise, Horn axioms are a special kind of Horn-$\mathcal{EL}$ ones, which are themselves a subclass of $\mathcal{EL}$ axioms. According to the semantics of these axioms, it is easy to see that dual-Horn KBs are not more expressive than core ones: a dual-Horn axiom $p \rightarrow q_1 \wedge \ldots \wedge q_n$ can be expressed by the core axioms $p \rightarrow q_1, \ldots, p \rightarrow q_n$. Hence, any dual-Horn KB can be transformed into an equivalent core KB in linear time. However, as we show in this paper, interestingly the complexity of pinpointing-related problems is in general higher for dual-Horn KBs than for core ones.

In complexity theory, we are sometimes interested not only in deciding whether a problem has a solution or not, but also in enumerating all solutions of the problem. We say that an algorithm runs with *polynomial delay* (Johnson, Yannakakis, and Papadimitriou 1988) if the time until the first solution is generated, and thereafter the time between any two consecutive solutions is bounded by a polynomial in the size of the input. We say that it runs in *output polynomial time* if it outputs all solutions in time polynomial in the size of the input *and the output*. In general, it is possible that an enumeration algorithm has exponentially many solutions. One advantage of an output polynomial algorithm is that it runs in polynomial time whenever the problem has polynomially many solutions. However, an output polynomial algorithm may for instance first compute all solutions and then output them all together. A polynomial delay algorithm on the other hand, outputs solutions with only polynomial time between them. This kind of algorithm is especially good if one wants to enumerate the solutions one at a time and maybe stop the execution before all of them have been found; for instance after k solutions have been output.

Complexity of Enumerating All MinAs

The main problem we consider here is, given a KB and a consequence of it, computing all MinAs for this consequence in the given KB. We start with the definition of a MinA.

Definition 1 (MinA). Let $\mathcal{K}$ be a set of axioms and φ be a logical consequence of it, i.e., $\mathcal{K} \models \varphi$. We call a set $\mathcal{M} \subseteq \mathcal{K}$ a *minimal axiom set* or *MinA* for φ in $\mathcal{K}$ if $\mathcal{M} \models \varphi$ and it is minimal w.r.t. set inclusion.

Our MinA enumeration problem is formally defined as follows:

Problem: MINA-ENUM
Input: A KB $\mathcal{K}$ and an axiom φ of the same type such that $\mathcal{K} \models \varphi$.
Output: The set of all MinAs for φ in $\mathcal{K}$.

Note that for core KBs, which are basically directed graphs, a MinA is a simple path between two given vertices, and enumerating all MinAs corresponds to enumerating all simple paths between two given vertices, which can easily be done with polynomial delay (Danielson 1968; Yen 1971). However, the situation is not so clear for Horn KBs. To the best of our knowledge, only (Nielsen, Pretolani, and Andersen 2006) considers a problem related to ours on directed hypergraphs, but it is not exactly the one considered here.

Enumeration without a Specific Order

We start with the Horn setting and show that for this kind of KBs MinAs can be efficiently enumerated by giving a polynomial delay algorithm. The algorithm depends on the notion of a valid ordering.

Definition 2 (Valid Ordering). Let $\mathcal{K}$ be a Horn KB, and $\phi = \bigwedge_{i=1}^{n} a_i \rightarrow b$ be an axiom in $\mathcal{K}$. We denote the left handside (lhs) of ϕ with $\mathsf{T}(\phi)$, and its right handside (rhs) with $\mathsf{h}(\phi)$, i.e., $\mathsf{T}(\phi) := \{a_1, \ldots, a_n\}$ and $\mathsf{h}(\phi) := b$. With $\mathsf{h}^{-1}(b)$ we denote the set of axioms in $\mathcal{K}$ whose rhs are b. Let $\mathcal{M} = \{t_1, \ldots, t_m\}$ be a MinA for $\bigwedge_{a \in \mathcal{A}} a \rightarrow c$. We call an ordering $t_1 < \ldots < t_m$ a *valid ordering* on $\mathcal{M}$ if for every $1 \leq i \leq m$, $\mathsf{T}(t_i) \subseteq \mathcal{A} \cup \{\mathsf{h}(t_1), \ldots, \mathsf{h}(t_{i-1})\}$ holds.[1]

It is easy to see that for every MinA $\mathcal{M}$ there is always at least one such valid ordering: the first elements are those having their lhs contained in $\mathcal{A}$, and later axioms are those that may contain also the rhs of previously included ones. In the following, we use this fact to construct from a given MinA a set of KBs that precisely contain the remaining MinAs.

[1] That is, each variable on the lhs of t_i is in $\mathcal{A}$, or it is the rhs of a previous axiom.

Algorithm 1 Enumerating all MinAs for Horn KBs

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1: ALL-MINAS(𝒦,ϕ)
2:                 ▷ (𝒦 a Horn KB, ϕ an axiom s.t. 𝒦 ⊨ ϕ)
3: if 𝒦 ⊭ ϕ then return
4: else
5:     ℳ := a MinA in 𝒦
6:     output ℳ
7:     for 1 ≤ i ≤ |ℳ| do
8:         compute 𝒦_i from ℳ as in Definition 3
9:         ALL-MINAS(𝒦_i,ϕ)
```

Definition 3 ($\mathcal{K}_i$). Let $\mathcal{M}$ be a MinA in $\mathcal{K}$ with $|\mathcal{M}| = m$, and $<$ be a valid ordering on $\mathcal{M}$. For each $1 \leq i \leq m$ we obtain a KB $\mathcal{K}_i$ from $\mathcal{K}$ as follows: (i) for each j s.t. $i < j \leq m$ remove all axioms in $\mathsf{h}^{-1}(\mathsf{h}(t_j))$ except for t_j, i.e., remove all axioms with the same rhs as t_j except for t_j itself; (ii) remove t_i.

The KBs constructed in Definition 3 represent a partition of all remaining MinAs in the sense that each MinA belongs to one, and only one, $\mathcal{K}_i$.

Lemma 4. *Let $\mathcal{M}$ be a MinA for ϕ in $\mathcal{K}$, and let $\mathcal{K}_1, \ldots, \mathcal{K}_m$ be constructed from $\mathcal{K}$ and $\mathcal{M}$ as in Definition 3. Then, for every MinA $\mathcal{N}$ for ϕ in $\mathcal{K}$ that is different from $\mathcal{M}$, there exists* **exactly one** *i, where $1 \leq i \leq m$, such that $\mathcal{N}$ is a MinA for ϕ in $\mathcal{K}_i$.*

Proof. Let $t_1 < \ldots < t_m$ be a valid ordering on $\mathcal{M}$, and $\mathcal{N}$ a MinA for ϕ in $\mathcal{K}$ such that $\mathcal{N} \neq \mathcal{M}$. Then, $\mathcal{M} \setminus \mathcal{N} \neq \emptyset$. Let t_k be the largest axiom in $\mathcal{M} \setminus \mathcal{N}$ w.r.t. the ordering $<$. We show that $\mathcal{N} \subseteq \mathcal{K}_k$ and $\mathcal{N} \not\subseteq \mathcal{K}_i$ for all $i \neq k$, $1 \leq i \leq m$.

Assume there is an axiom $t \in \mathcal{N}$ s.t. $t \notin \mathcal{K}_k$. t should be one of the axioms removed from $\mathcal{K}$ either in step (i), or in step (ii) of Definition 3. It cannot be step (ii) because $t_k \notin \mathcal{N}$ since $t_k \in \mathcal{M} \setminus \mathcal{N}$. Thus it should be step (i). This implies that there exists a j, $k < j \leq m$, such that t_j satisfies $\mathsf{h}(t) = \mathsf{h}(t_j)$. Recall that we chose j to be the largest axiom in $\mathcal{M} \setminus \mathcal{N}$ w.r.t. the valid ordering $<$ on $\mathcal{M}$. Then this t_j should be in $\mathcal{N}$. But then $\mathcal{N}$ contains two axioms with the rhs $\mathsf{h}(t)$, which contradicts with the fact that $\mathcal{N}$ is a MinA, and thus it is minimal. Hence, $\mathcal{N} \subseteq \mathcal{K}_k$.

Now take an i s.t. $i \neq k$. If $i > k$, then $t_i \in \mathcal{N}$ but $t_i \notin \mathcal{K}_i$, and hence $\mathcal{N} \not\subseteq \mathcal{K}_i$. If $i < k$, then there is an axiom $t \in \mathcal{N}$ such that $\mathsf{h}(t) = \mathsf{h}(t_k)$ since otherwise $\mathcal{M}$ and $\mathcal{N}$ would not be MinAs. By construction, $t \notin \mathcal{K}_i$, hence $\mathcal{N} \not\subseteq \mathcal{K}_i$. $\square$

Lemma 4 gives an idea of how to compute the remaining MinAs from a given one. Algorithm 1 describes a procedure that uses this lemma for the enumeration of all MinAs.

Theorem 5. *Algorithm 1 solves* MINA-ENUM *for Horn KBs with polynomial delay.*

Proof. The algorithm terminates since $\mathcal{K}$ is finite. It is sound since its outputs are MinAs for ϕ in $\mathcal{K}$. Completeness follows from Lemma 4.

In each recursive call of the algorithm there is one consequence check (line 3), and one MinA computation (line 5).

The consequence check can be done in polynomial time by the well-known linear-time algorithm in (Dowling and Gallier 1984). One MinA can be computed in polynomial time by iterating over the axioms in $\mathcal{K}$ and removing an axiom if remaining ones still have the consequence. Thus the algorithm spends at most polynomial time between each output, i.e., it is polynomial delay. □

Next we consider MINA-ENUM for dual-Horn KBs. For this, we first investigate the following decision problem which is is closely related to MINA-ENUM. As we will see, determining its complexity is important for determining the complexity of MINA-ENUM.

Problem: ALL-MINAS
Input: A KB $\mathcal{K}$ and an axiom φ of the same type such that $\mathcal{K} \models \varphi$, and a set of KBs $\mathscr{K} \subseteq \mathscr{P}(\mathcal{K})$.
Question: Is $\mathscr{K}$ precisely the set of all MinAs for φ in $\mathcal{K}$?

As Proposition 6 shows, if ALL-MINAS cannot be decided in polynomial time, then MINA-ENUM cannot be solved in output polynomial time. Its proof is based on a generic argument, which can also be found in (Eiter and Gottlob 1995b) Theorem 4.5, but for the sake of completeness we present it here once more.

Proposition 6. *If* ALL-MINAS *cannot be decided in polynomial time, then* MINA-ENUM *cannot be solved in output-polynomial time.*

Proof. Assume we have an algorithm A that solves MINA-ENUM in output-polynomial time. Let its runtime be bounded by a polynomial $p(IS, OS)$ where IS denotes the size of the input KB and OS denotes the size of the output, i.e., the set of all MinAs.

In order to decide ALL-MINAS for an instance given by $\mathcal{K}$, φ, and $\mathscr{K} \subseteq \mathscr{P}(\mathcal{K})$, we construct another algorithm A' that works as follows: it runs A on $\mathcal{K}$ and φ for at most $p(|\mathcal{K}|, |\mathscr{K}|)$-many steps. If A terminates within this many steps, then A' compares the output of A with $\mathscr{K}$ and returns *yes* if and only if they are equal. If they are not equal, A' returns *no*. If A has not yet terminated after $p(|\mathcal{K}|, |\mathscr{K}|)$-many steps, this implies that there is at least one MinA that is not contained in $\mathscr{K}$, so A' returns *no*. It is easy to see that the runtime of A' is bounded by a polynomial in $|\mathcal{K}|$ and $|\mathscr{K}|$, that is A' decides ALL-MINAS in polynomial time. □

The proposition shows that the complexity of ALL-MINAS is indeed closely related to the complexity of MINA-ENUM. We now present some hardness results for enumerating MinAs when other types of KBs different from Horn are used. It is not difficult to see that, for all types of axioms considered in this paper, ALL-MINAS is in coNP: given an instance of ALL-MINAS, a nondeterministic algorithm can guess a subset of $\mathcal{K}$ that is not in $\mathscr{K}$, and in polynomial time verify that this is a MinA, thus $\mathscr{K}$ is *not* the set of all MinAs. In the following we show that for dual-Horn KBs ALL-MINAS is at least as hard as recognizing the set of all minimal transversals of a given hypergraph. However, whether it is coNP-hard remains unfortunately open. We later show that ALL-MINAS is coNP-complete if Horn-$\mathcal{EL}$ axioms are considered.

First we briefly recall some basic notions on hypergraphs. A *hypergraph* $\mathcal{H} = (V, \mathcal{E})$ consists of a set of *vertices* $V = \{v_i \mid 1 \leq i \leq n\}$, and a set of *(hyper)edges* $\mathcal{E} = \{E_j \mid 1 \leq j \leq m\}$ where $E_j \subseteq V$. We assume w.l.o.g. that the set of edges as well as the set of vertices is nonempty, and the union of all edges yields the vertex set. A set $W \subseteq V$ is called a *transversal* of $\mathcal{H}$ if it intersects all edges of $\mathcal{H}$, i.e., $\forall E \in \mathcal{E}.\ E \cap W \neq \emptyset$. A transversal is called *minimal* if no proper subset of it is a transversal. The set of all minimal transversals of $\mathcal{H}$ constitutes another hypergraph on V called the *transversal hypergraph* of $\mathcal{H}$, which is denoted by $Tr(\mathcal{H})$. Generating $Tr(\mathcal{H})$ is an important problem which has applications in many fields of computer science (Hagen 2008). The well-known decision problem associated to this computation problem is defined as follows:

Problem: TRANSVERSAL HYPERGRAPH (TRANS-HYP)
Input: Two hypergraphs $\mathcal{H} = (V, \mathcal{E}_\mathcal{H})$ and $\mathcal{G} = (V, \mathcal{E}_\mathcal{G})$.
Question: Is $\mathcal{G}$ the transversal hypergraph of $\mathcal{H}$, i.e., does $Tr(\mathcal{H}) = \mathcal{G}$ hold?

TRANS-HYP is known to be in coNP, but its lower bound is a prominent open problem. More precisely, so far neither a polynomial time algorithm has been found, nor has it been proved to be coNP-hard. In a landmark paper (1996) Fredman and Khachiyan proved that TRANS-HYP can be solved in $n^{o(\log n)}$ time, which implies that this problem is most likely not coNP-hard. It is conjectured that this problem, together with several computationally equivalent problems, forms a class properly contained between P and coNP (Fredman and Khachiyan 1996).

Theorem 7. ALL-MINAS *is* TRANS-HYP*-hard for dual-Horn KBs.*

Proof. Let an instance of TRANS-HYP be given by the hypergraphs $\mathcal{H} = (V, \mathcal{E}_\mathcal{H})$ and $\mathcal{G} = (V, \mathcal{E}_\mathcal{G})$. From $\mathcal{H}$ and $\mathcal{G}$ we construct an instance of ALL-MINAS as follows: for every vertex $v \in V$ we introduce a propositional variable p_v, for every edge $E \in \mathcal{E}_\mathcal{H}$ a propositional variable p_E, and finally one additional propositional variable a. For constructing a dual-Horn KB from $\mathcal{H}$ and a set of vertices $W \subseteq V$, we define the following operator, which is also going to be used in later proofs:

$$\mathcal{K}_{W,\mathcal{H}} := \{p_v \rightarrow \bigwedge_{v \in E, E \in \mathcal{E}_\mathcal{H}} p_E \mid v \in W\} \cup \{a \rightarrow \bigwedge_{v \in V} p_v\}.$$

Using these we construct the KB $\mathcal{K} := \mathcal{K}_{V,\mathcal{H}}$, a set of KBs $\mathscr{K} := \{\mathcal{K}_{E,\mathcal{H}} \mid E \in \mathcal{E}_\mathcal{G}\} \subseteq \mathscr{P}(\mathcal{K})$, and the axiom $\varphi := a \rightarrow \bigwedge_{E \in \mathcal{E}_\mathcal{H}} p_E$ that follows from $\mathcal{K}$. Obviously this construction creates an instance of ALL-MINAS for dual-Horn KBs and it can be done in time polynomial in the sizes of $\mathcal{H}$ and $\mathcal{G}$.

We claim that $\mathcal{G}$ is the transversal hypergraph of $\mathcal{H}$ if and only if $\mathscr{K}$ is precisely the set of all MinAs for φ in $\mathcal{K}$. Note that $a \rightarrow \bigwedge_{v \in V} p_v$ is the only axiom in $\mathcal{K}$ such that a appears on the lhs, which implies that every MinA must contain this axiom. Hence, every MinA is of the form $\mathcal{K}_{W,\mathcal{H}}$ for some $W \subseteq V$. To prove our claim, it suffices to show that a set of vertices $W \subseteq V$ is a minimal transversal of $\mathcal{H}$ if and only if the set of axioms $\mathcal{K}_{W,\mathcal{H}}$ is a MinA.

($\Rightarrow$) Assume that W is a minimal transversal of $\mathcal{H}$. By definition W satisfies $W \cap E \neq \emptyset$ for every $E \in \mathcal{E}_H$. This implies that $\mathcal{K}_{W,\mathcal{H}} \models \varphi$ holds. Moreover, $\mathcal{K}_{W,\mathcal{H}}$ is minimal since W is minimal, i.e., $\mathcal{K}_{W,\mathcal{H}}$ is a MinA.

($\Leftarrow$) Now assume that $\mathcal{K}_{W,\mathcal{H}}$ is a MinA. Then every p_E where $E \in \mathcal{E}_\mathcal{H}$ appears on the rhs of at least one of the axioms in $\mathcal{K}_{W,\mathcal{H}}$. This implies that W intersects every E, i.e., it is a transversal of $\mathcal{H}$. Moreover it is minimal since $\mathcal{K}_{W,\mathcal{H}}$ is minimal. □

A direct consequence of this theorem is that the enumeration of all MinAs in a dual-Horn KB is at least as hard as the enumeration of the transversals of a hypergraph.

Corollary 8. MINA-ENUM *for dual-Horn KBs is at least as hard as enumerating hypergraph transversals.*

Up to now we have investigated the complexity of MINA-ENUM for the propositional case. In particular, we have presented a polynomial delay algorithm for enumerating all MinAs in a Horn KB. However, whether such an algorithm exists for dual-Horn KBs remains open. We now turn our attention to $\mathcal{EL}$ KBs, and show that there is no output polynomial algorithm that enumerates all MinAs in a Horn-$\mathcal{EL}$ KB, unless P $=$ NP. As a first step to this result, we show that ALL-MINAS is intractable for Horn-$\mathcal{EL}$ KBs.

Theorem 9. ALL-MINAS *is co*NP*-complete for Horn-$\mathcal{EL}$ TBoxes.*

Proof. We have already shown that it is in coNP. To show coNP-hardness, we present a reduction from the following coNP-hard problem (Eiter and Gottlob 1991; Baader, Peñaloza, and Suntisrivaraporn 2007).

Problem: ALL-MV
Input: A monotone Boolean formula ϕ and a set $\mathscr{V}$ of minimal valuations satisfying ϕ.
Question: Is $\mathscr{V}$ precisely the set of all minimal valuations satisfying ϕ?

Let $\phi, \mathscr{V}$ be an instance of ALL-MV; we denote as $\mathsf{sub}(\phi)$ the set of all subformulas of ϕ, and define $\mathsf{csub}(\phi) := \mathsf{sub}(\phi) \setminus \{p \in \mathsf{sub}(\phi) \mid p \text{ is a propositional variable}\}$. We introduce three concept names B_ψ, C_ψ, D_ψ, and two role names r_ψ, s_ψ for every subformula ψ of ϕ and two additional concept names A and E. For each $\psi \in \mathsf{sub}(\phi)$ we define a TBox $\mathcal{T}_\psi$ as follows: if ψ is the propositional variable p, then $\mathcal{T}_\psi := \{A \sqsubseteq B_p\}$; if $\psi = \psi_1 \wedge \psi_2$, then $\mathcal{T}_\psi := \{A \sqsubseteq \exists r_\psi.C_\psi, C_\psi \sqsubseteq B_{\psi_1}, C_\psi \sqsubseteq B_{\psi_2}, \exists r_\psi.B_\psi \sqsubseteq D_\psi, B_{\psi_1} \sqcap B_{\psi_2} \sqsubseteq B_\psi\}$; if $\psi = \psi_1 \vee \psi_2$, then $\mathcal{T}_\psi := \{A \sqsubseteq \exists r_\psi.B_{\psi_1}, A \sqsubseteq \exists s_\psi.B_{\psi_2}, \exists r_\psi.B_\psi \sqcap \exists s_\psi.B_\psi \sqsubseteq D_\psi, B_{\psi_1} \sqsubseteq B_\psi, B_{\psi_2} \sqsubseteq B_\psi\}$. Finally, we set

$$\mathcal{T} := \bigcup_{\psi \in \mathsf{sub}(\phi)} \mathcal{T}_\psi \cup \{ \bigsqcap_{\psi \in \mathsf{csub}(\phi)} D_\psi \sqcap B_\phi \sqsubseteq E\}.$$

Notice that for every $\mathcal{T}' \subseteq \mathcal{T}$, if $\mathcal{T}' \models A \sqsubseteq E$, then also $A \sqsubseteq D_\psi$ for every $\psi \in \mathsf{csub}(\phi)$. But in order to have $A \sqsubseteq D_\psi$, all the axioms in $\mathcal{T}_\psi$ are necessary, and thus $\mathcal{T}_\psi \subseteq \mathcal{T}'$. In particular, if $\psi = \psi_1 \wedge \psi_2$, then $B_{\psi_1} \sqcap B_{\psi_2} \sqsubseteq B_\psi \in \mathcal{T}'$, and if $\psi = \psi_1 \vee \psi_2$, then $\{B_{\psi_1} \sqsubseteq B_\psi, B_{\psi_2} \sqsubseteq B_\psi\} \subseteq \mathcal{T}'$. Thus, a valuation $\mathcal{V}$ satisfies ϕ iff $\mathcal{T}_\mathcal{V} := \{A \sqsubseteq B_p \mid p \in \mathcal{V}\} \cup \bigcup_{\psi \in \mathsf{csub}(\phi)} \mathcal{T}_\psi \cup \{\bigsqcap_{\psi \in \mathsf{csub}(\phi)} D_\psi \sqcap B_\phi \sqsubseteq E\}$ entails $A \sqsubseteq E$. This in particular shows that $\mathscr{V}$ is the set of all minimal valuations satisfying ϕ iff $\{\mathcal{T}_\mathcal{V} \mid \mathcal{V} \in \mathscr{V}\}$ is the set of all MinAs for $A \sqsubseteq E$ in $\mathcal{T}$. □

The following is an immediate consequence of Theorem 9 and Proposition 6.

Corollary 10. *For Horn-$\mathcal{EL}$ TBoxes* MINA-ENUM *cannot be solved in output polynomial time, unless* P $=$NP.

Enumeration in a Specified Order

We now consider the case when MinAs are required to be output in a specified lexicographic order. The lexicographic order we use is defined as follows:

Definition 11 (Lexicographic Order). Let the elements of a set S be linearly ordered. This order induces a linear strict order on $\mathscr{P}(S)$, which is called the *lexicographic order*. We say that a set $R \subseteq S$ is *lexicographically smaller* than a set $T \subseteq S$ where $R \neq T$ if the first element at which they disagree is in R.

We first look at the complexity of finding the lexicographically first MinA.

Problem: FIRST-MINA
Input: A KB $\mathcal{K}$ and an axiom φ of the same type such that $\mathcal{K} \models \varphi$, a MinA $\mathcal{M}$ for φ in $\mathcal{K}$, and a linear order on $\mathcal{K}$.
Question: Is $\mathcal{M}$ the first MinA w.r.t. the lexicographic order induced by the given linear order?

This problem is of particular interest when for instance one can assign a degree of trust to the axioms in the KB. In this setting if we order the axioms in such a way that less trusted axioms appear before the more trusted ones, the lexicographically first MinA will be the one that has the most distrusted axioms, and hence the most likely cause of an error. As we show now, finding the lexicographically first MinA is coNP-complete for dual-Horn and Horn-$\mathcal{EL}$ KBs.

Theorem 12. FIRST-MINA *is co*NP*-complete for dual-Horn KBs.*

Proof. The problem is in coNP. If $\mathcal{M}$ is not the lexicographically first MinA, a proof of this can be given by guessing a subset of $\mathcal{K}$ and verifying in polynomial time that it is a MinA, and it is lexicographically smaller than $\mathcal{M}$.

In order to show coNP-hardness, we present a reduction from the problem of checking whether a given maximal independent set is lexicographically the last maximal independent set of a given graph. Recall that a *maximal independent set* of a graph $\mathcal{G} = (V, \mathcal{E})$ is a subset $V' \subseteq V$ of the vertices such that no two vertices in V' are joined by an edge in $\mathcal{E}$, and each vertex in $V \setminus V'$ is joined by an edge to some vertex in V'. This problem is known to be coNP-complete (Johnson, Yannakakis, and Papadimitriou 1988).

Problem: LAST MAX. INDEPENDENT SET (LAST-MIS)
Input: A graph $\mathcal{G} = (V, \mathcal{E})$, a maximal independent set $S \subseteq V$, and a linear order on V.

Question: Is S the last maximal independent set w.r.t. the lexicographic order induced by the given linear order?

Let an instance of LAST-MIS be given with the graph $\mathcal{G} = (V, \mathcal{E})$ and the maximal independent set S. From $\mathcal{G}$ and S we construct an instance of FIRST-MINA as follows: we construct the sets $\mathcal{K}_{W,\mathcal{G}}$ as in the proof of Theorem 7, and consider the axiom $\varphi := a \rightarrow \bigwedge_{E \in \mathcal{E}} p_E$ that follows from $\mathcal{K}_{V,\mathcal{G}}$. Additionally by using S we construct the set of axioms $\mathcal{M} := \mathcal{K}_{V \setminus S,\mathcal{G}}$. Note that $\mathcal{K}_{V,\mathcal{G}}$ contains exactly $|V|+1$ axioms. We order these axioms such that an axiom with premise p_v comes before the axiom with premise $p_{v'}$ if and only if the vertex v comes before the vertex v' in the originally given linear order on V. Finally we place φ as the last one. It is easy to see that this construction indeed creates an instance of FIRST-MINA for dual-Horn KBs, and it can be done in time polynomial in the sizes of $\mathcal{G}$ and S. We claim that S is lexicographically the last maximal independent set if and only if $\mathcal{M}$ is lexicographically the first MinA.

($\Rightarrow$) Assume S is lexicographically the last maximal independent set. Then $V \setminus S$ contains at least one vertex from every edge (i.e., it is a vertex cover), since otherwise S would not be an independent set. Thus every p_E, for $E \in \mathcal{E}$, appears on the rhs of at least one axiom in $\mathcal{M}$. That is $\mathcal{M} \models \varphi$ holds. Since S is maximal, $V \setminus S$ and thus $\mathcal{M}$ is minimal, i.e., $\mathcal{M}$ is a MinA. Moreover it is lexicographically the first one since S is lexicographically the last maximal independent set.

($\Leftarrow$) Assume $\mathcal{M}$ is lexicographically the first MinA. Then every p_E, for $E \in \mathcal{E}$, appears on the rhs of at least one axiom in $\mathcal{M}$ since otherwise $\mathcal{M} \models \varphi$ would not hold. That is, $V \setminus S$ contains at least one vertex from every edge. Then S contains at most one vertex from every edge, i.e., it is an independent set. Since $\mathcal{M}$ is minimal, $V \setminus S$ is also minimal, and thus S is maximal. That is, S is a maximal independent set. Moreover it is lexicographically the last one since $\mathcal{M}$ is lexicographically the first MinA. □

Since generating the lexicographically first MinA is already intractable, Theorem 12 has the following consequence:

Corollary 13. *Unless* P = NP*, MinAs cannot be enumerated for dual-Horn KBs in lexicographic order with polynomial delay.*

Next we consider the same problem for Horn-$\mathcal{EL}$ KBs. As done in the dual-Horn case, we show first that FIRST-MINA is coNP-complete, and use that result to show that it is not possible to enumerate all MinAs for Horn-$\mathcal{EL}$ KBs in polynomial delay, unless P = NP.

Theorem 14. FIRST-MINA *is co*NP*-complete for Horn-$\mathcal{EL}$ KBs.*

Proof. The problem is clearly in coNP. To show hardness, we give a reduction from LAST-MIS. Let $\mathcal{G} = (V, \mathcal{E})$ and S be an instance of LAST-MIS. From $\mathcal{G}$ we construct a Horn-$\mathcal{EL}$ TBox $\mathcal{T}$ as follows: first we introduce a concept P_E for every $E \in \mathcal{E}$, and concepts P_v, Q_v and role name r_v for each $v \in V$, and additionally two concept names A, B. For every $v \in V$ we construct the TBox $\mathcal{T}_v := \{P_v \sqsubseteq P_E \mid v \in E, E \in \mathcal{E}_\mathcal{G}\} \cup \{A \sqsubseteq \exists r_v.P_v, \bigsqcap_{v \in E, E \in \mathcal{E}_\mathcal{G}} \exists r_v.P_E \sqsubseteq Q_v\}$. We then define the set $\mathcal{T}_f := \bigcup_{v \in V} \mathcal{T}_v \cup \{\bigsqcap_{E \in \mathcal{E}_\mathcal{H}} P_E \sqcap \bigsqcap_{v \in V} Q_v \sqsubseteq B\}$, and finally, for a set of $W \subseteq V$, we define $\mathcal{T}_W := \mathcal{T}_f \cup \{A \sqsubseteq P_v \mid v \in W\}$.

Notice that for every $\mathcal{T}' \subseteq \mathcal{T}$, if $\mathcal{T}' \models A \sqsubseteq Q_v$, then $\mathcal{T}_v \subseteq \mathcal{T}'$. Hence, if $\mathcal{T}' \models A \sqsubseteq B$, then $\mathcal{T}_f \subseteq \mathcal{T}'$. Furthermore, $S \subseteq V$ is an independent set iff $\mathcal{T}_{V \setminus S} \models A \sqsubseteq B$.

We now order the axioms in $\mathcal{T}_V$ as follows: first appear all the axioms $A \sqsubseteq P_v$ using the same order of V, and afterwards are all the axioms in $\mathcal{T}_f$ in any order. Then S is the last maximal independent set iff $\mathcal{T}_{V \setminus S}$ is the first MinA for $A \sqsubseteq B$ in $\mathcal{T}_V$. □

Algorithm 2 Enumerating all MinAs in reverse lex. order

ALL-MINAS-REV-ORDER($\mathcal{K}$,ϕ)
  ▷ ($\mathcal{K}$ a Horn KB, ϕ an axiom s.t. $\mathcal{K} \models \phi$)
$\mathcal{Q} := \{\mathcal{K}\}$
while $\mathcal{Q} \neq \emptyset$ **do**
  $\mathcal{J}$:= maximum element of $\mathcal{Q}$
  remove $\mathcal{J}$ from $\mathcal{Q}$
  $\mathcal{M}$:= the lex. largest MinA in $\mathcal{J}$
  output $\mathcal{M}$
  for $1 \leq i \leq |\mathcal{M}|$ **do**
    compute $\mathcal{K}_i$ from $\mathcal{M}$ as in Definition 3
    insert $\mathcal{K}_i$ into $\mathcal{Q}$ if $\mathcal{K}_i \models \phi$

Although computing the first MinA is coNP-hard for both dual-Horn and Horn-$\mathcal{EL}$ KBs, interestingly computing the last MinA is polynomial for all types of KBs we consider here. We start iterating over the axioms of the KB with the axiom that is the smallest one w.r.t. the linear order on the KB, and remove an axiom if the remaining ones still have the given conseqence. It is easy to see that the resulting set of axioms is lexicographically the last MinA. Even more interestingly, we now give an algorithm for Horn KBs that enumerates MinAs in reverse lexicographic order with polynomial delay.

Our algorithm keeps a set of KBs in a priority queue $\mathcal{Q}$. These KBs are the "candidates" from which the MinAs are going to be computed. Each KB can contain zero or more MinAs. They are inserted into $\mathcal{Q}$ by the algorithm at a cost of $O(n \cdot log(M))$ per insertion, where n is the size of the original KB and M is the total number of such KBs inserted. Note that M can be exponentially larger than n since there can be exponentially many MinAs. That is, the algorithm uses potentially exponential space. The other operation that the algorithm performs on $\mathcal{Q}$ is to find and delete the maximum element of $\mathcal{Q}$. The maximum element of $\mathcal{Q}$ is the KB in $\mathcal{Q}$ that contains the lexicographically largest MinA among the MinAs contained in all other KBs in $\mathcal{Q}$. This operation can also be performed within a $O(n \cdot log(M))$ time bound. The time bounds for insertion and deletion depend also on n since they require a last MinA computation.

Theorem 15. *Algorithm 2 enumerates MinAs in the Horn setting in reverse lexicographic order with polynomial delay.*

Proof. The algorithm terminates since $\mathcal{K}$ is finite. Soundness is shown as follows: $\mathcal{Q}$ contains initially only the original KB $\mathcal{K}$. Thus the first output is lexicographically the last MinA in $\mathcal{K}$. By Lemma 4 the MinA that comes just before the last one is contained in exactly one of the $\mathcal{K}_i$s that are computed and inserted into $\mathcal{Q}$ in lines 10 and 11. In line 5 $\mathcal{J}$ is assigned the KB that contains this MinA. Thus the next output will be the MinA that comes just before lexicographically the last one. It is not difficult to see that in this way the MinAs will be enumerated in reverse lexicographic order. By Lemma 4 it is guaranteed that the algorithm enumerates all MinAs.

In one iteration, the algorithm performs one find operation and one delete operation on $\mathcal{Q}$, which both take time $O(n \cdot log(M))$, and a MinA computation that takes $O(n)$ time. In addition it performs at most n $\mathcal{K}_i$ computations, and at most n insertions into $\mathcal{Q}$. Each $\mathcal{K}_i$ computation takes $O(n^2)$ time, and each insertion takes $O(n \cdot log(M))$ time. The total delay is thus $O(2 \cdot (n \cdot log(M)) + n + n \cdot (n^2 + n \cdot log(M))) = O(n^3)$. $\square$

Preferred and Unwanted Axioms

Next we investigate the problem of existence of a MinA that does not subsume any of the given sets of axioms. This problem can be useful in applications where one wants to avoid certain combinations of axioms in the MinAs.

Problem: MINA-IRRELEVANCE
Input: A KB $\mathcal{K}$ and an axiom φ of the same type such that $\mathcal{K} \models \varphi$, and a set $\mathscr{K} \subseteq \mathscr{P}(\mathcal{K})$.
Question: Is there a MinA $\mathcal{M}$ for φ in $\mathcal{K}$ such that $\mathcal{S} \nsubseteq \mathcal{M}$ for every $\mathcal{S} \in \mathscr{K}$?

MINA-IRRELEVANCE refers to the problem of deciding whether there is a MinA that does not contain any of the sets in $\mathscr{K}$. Intuitively, one can consider $\mathscr{K}$ as a collection of sets of axioms that are already known to be faulty. Hence, any MinA that is a superset of any element of $\mathscr{K}$ will give no further information about the causes of an erroneous consequence. In order to decide MINA-IRRELEVANCE, it does not suffice to remove the axioms that appear in one or all the sets that form $\mathscr{K}$. Indeed, there can still be a MinA that has a non-empty intersection with each element of $\mathscr{K}$, but is not a superset of any of them. The most direct approach for solving MINA-IRRELEVANCE is to test for each hitting set $\mathcal{S}$ of $\mathscr{K}$, whether there is a MinA that does not contain any of the axioms in $\mathcal{S}$.[2] However, there can be exponentially many such hitting sets in the size of $\mathscr{K}$, which means that this simple approach cannot avoid an exponential execution time in the worst case. We now show that the problem is in fact NP-complete for dual-Horn and Horn-$\mathcal{EL}$ KBs.

Theorem 16. MINA-IRRELEVANCE *is* NP*-complete for dual-Horn KBs.*

Proof. The problem is clearly in NP. A nondeterministic algorithm for solving it first guesses a set $\mathcal{M} \subseteq \mathcal{K}$, then tests in polynomial time whether it is a MinA that does not contain any of the $\mathcal{S}$ in $\mathscr{K}$. For showing hardness we give a reduction from the NP-hard hypergraph 2-coloring problem (Garey and Johnson 1990).

Problem: HYPERGRAPH 2-COLORING
Input: A hypergraph $\mathcal{H} = (V, \mathcal{E})$.
Question: Is $\mathcal{H}$ 2-colorable, i.e., is there a $W \subseteq V$ such that for all $E \in \mathcal{E}$, $W \cap E \neq \emptyset$ and $(V \setminus W) \cap E \neq \emptyset$?

Let an instance of HYPERGRAPH 2-COLORING be given with the hypergraph $\mathcal{H} = (V, \mathcal{E})$. We construct an instance of MINA-IRRELEVANCE as follows: as in the proof of Theorem 7, we construct the KB $\mathcal{K} := \mathcal{K}_{V,\mathcal{H}}$ and the axiom φ contructed there, as well as a set of KBs $\mathscr{K} = \{\mathcal{K}_{E,\mathcal{H}} \mid E \in \mathcal{E}\}$. It is easy to see that this construction indeed creates an instance of MINA-IRRELEVANCE for dual-Horn KBs and it can be done in time polynomial in the size of $\mathcal{H}$. We claim that $\mathcal{H}$ is 2-colorable if and only if there is a MinA $\mathcal{M}$ for φ in $\mathcal{K}$ such that $\mathcal{M}$ satisfies $\mathcal{S} \nsubseteq \mathcal{M}$ for every $\mathcal{S} \in \mathscr{K}$.

($\Rightarrow$) Assume $\mathcal{H}$ is 2-colorable. Then there is a $W \subseteq V$ such that $W \cap E \neq \emptyset$ and $(V \setminus W) \cap E \neq \emptyset$ for every $E \in \mathcal{E}$, i.e., both W and its complement are transversals of $\mathcal{H}$. Assume w.l.o.g. that W is minimal. We claim that $\mathcal{K}_{W,\mathcal{H}}$ is the MinA we are looking for. Since W is a transversal, every p_E for $E \in \mathcal{E}$, appears on the rhs of at least one axiom in $\mathcal{K}_{W,\mathcal{H}}$. That is $\mathcal{K}_{W,\mathcal{H}} \models \varphi$ holds. $\mathcal{K}_{W,\mathcal{H}}$ is minimal since W is minimal. Moreover, since $V \setminus W$ is also a transversal, every edge $E \in \mathcal{E}$ contains at least one vertex that is *not* in W. Thus every $\mathcal{S} \in \mathscr{K}$ contains at least one axiom that is *not* in $\mathcal{K}_{W,\mathcal{H}}$. In other words, $\mathcal{K}_{W,\mathcal{H}}$ is a MinA that is not a superset of any $\mathcal{S} \in \mathscr{K}$.

($\Leftarrow$) Assume $\mathcal{M}$ is a MinA that is not a superset of any $\mathcal{S} \in \mathscr{K}$. Define the set $W_{\mathcal{M}} = \{v \mid p_v \rightarrow \bigwedge_{v \in E, E \in \mathcal{E}} p_E \in \mathcal{M}\}$. Since $\mathcal{M}$ is a MinA for φ, for every $E \in \mathcal{E}$ it contains at least one axiom on whose rhs p_E occurs. That is, $W_{\mathcal{M}}$ intersects every $E \in \mathcal{E}$. Since $\mathcal{M}$ is not a superset of any $\mathcal{S} \in \mathscr{K}$, every $\mathcal{S}$ contains at least one axiom that is *not* in $\mathcal{M}$. This that every $E \in \mathcal{E}$ contains at least one vertex that is *not* in $W_{\mathcal{M}}$. That is, $V \setminus W_{\mathcal{M}}$ intersects every $E \in \mathcal{E}$. Thus we have shown that $W_{\mathcal{M}}$ is a 2-coloring of $\mathcal{H}$. $\square$

Next we show that for Horn-$\mathcal{EL}$ KBs the problem is NP-complete as well.

Theorem 17. MINA-IRRELEVANCE *is* NP*-complete for Horn-$\mathcal{EL}$ KBs.*

Proof. The problem is clearly in NP. We show NP-hardness by a reduction from the HYPERGRAPH 2-COLORING problem. Let $\mathcal{H} = (V, \mathcal{E})$ be a hypergraph; we construct the TBoxes $\mathcal{T}_v, \mathcal{T}_f$ and $\mathcal{T}_V$ as in the proof of Theorem 14. It is easy to see that $\mathcal{T} := \mathcal{T}_V, \phi := A \sqsubseteq B$ and the set of TBoxes $\mathscr{K} := \{\mathcal{T}_E \mid E \in \mathcal{E}\}$ form an instance of MINA-IRRELEVANCE for Horn-$\mathcal{EL}$ TBoxes. Furthermore, we know that for every $W \subseteq V$, W is a transversal of $\mathcal{H}$ iff $\mathcal{T}_W$ is a MinA for ϕ in $\mathcal{T}$. The hypergraph $\mathcal{H}$ is 2-colorable iff there is a transversal W of $\mathcal{H}$ such that for all $E \in \mathcal{E}, E \nsubseteq W$. Hence, $\mathcal{H}$ is 2-colorable iff there is a MinA $\mathcal{T}'$ for ϕ in $\mathcal{T}$ such that $\mathcal{T}_E \nsubseteq \mathcal{T}'$ for all $E \in \mathcal{E}$. $\square$

Next we consider the dual problem, which consists of checking the existence of a MinA that contains a certain axiom.

[2] Given a collection of sets $\mathscr{K}$, a *hitting set for* $\mathscr{K}$ is a set $\mathcal{S}$ that satisfies $\mathcal{S} \cap \mathcal{K} \neq \emptyset$ for every $\mathcal{K} \in \mathscr{K}$.

	FIRST-MINA	LAST-MINA	ALL-MINAS	MINA-REL	MINA-IRREL	MINA-ENUM		
						in lexicographic order		unordered
						forward	backward	
core		poly	poly			output poly	poly delay	poly delay
Horn		poly	poly	NP-c		output poly	poly delay	poly delay
dual-Horn	coNP-c	poly	TH-h		NP-c	not poly delay	TE-h	TE-h
Bool	coNP-c	poly	TH-h	NP-c	NP-c	not poly delay	TE-h	TE-h
Horn-$\mathcal{EL}$	coNP-c	poly	coNP-c	NP-c	NP-c	not output poly		

Table 2: Summary of the results

Problem: MINA-RELEVANCE
Input: A KB $\mathcal{K}$ and an axiom φ of the same type such that $\mathcal{K} \models \varphi$, and an axiom $\psi \in \mathcal{K}$.
Question: Is there a MinA $\mathcal{M}$ for φ in $\mathcal{K}$ such that $\psi \in \mathcal{M}$?

If we identify a specific axiom ψ as a possible culprit for an erroneous consequence from a KB, MINA-RELEVANCE would allow us to decide whether ψ indeed appears in at least one MinA, and hence influences the deduction of the consequence from the KB. We now show that this problem is NP-complete for Horn KBs.

Theorem 18. MINA-RELEVANCE *is* NP*-complete for Horn KBs.*

Proof. The problem is clearly in NP. A nondeterministic algorithm for solving it first guesses a subset of $\mathcal{K}$, then tests in polynomial time whether it is a MinA containing ψ. For showing hardness we are going to give a reduction from the following NP-complete problem (Eiter and Gottlob 1995a):

Problem: HORN-RELEVANCE
Input: Two sets of propositional variables H and M, a set $\mathcal{C}$ of definite Horn clauses over $H \cup M$, and a propositional variable $p \in H$.
Question: Is there a minimal $G \subseteq H$ such that $G \cup \mathcal{C} \models M$ and $p \in G$?

Let an instance of HORN-RELEVANCE be given with $H, M, \mathcal{C}$ and p. We construct an instance of MINA-RELEVANCE as follows: In addition to the propositional variables in $H \cup M$, we introduce two more fresh ones a, and b. Using these we construct the Horn KB $\mathcal{K} := \{a \to h \mid h \in H\} \cup \mathcal{C} \cup \{\bigwedge_{m \in M} m \to b\}$, the axiom $\varphi := a \to b$, and the axiom $\psi := a \to p$. It is easy to see that this construction indeed creates an instance of MINA-RELEVANCE and it can be done in polynomial time. We claim that there is a minimal $G \subseteq H$ such that $G \cup \mathcal{C} \models M$ and $p \in G$ if and only if there is a MinA $\mathcal{M}$ for φ in $\mathcal{K}$ such that $\psi \in \mathcal{M}$.

($\Rightarrow$) Assume that there is such a minimal G. From G we construct $\mathcal{K}_G := \{a \to g \mid g \in G\} \cup \mathcal{C} \cup \{\bigwedge_{m \in M} m \to b\}$. $\mathcal{K}_G \models a \to b$ since $G \cup \mathcal{C} \models M$. Thus, there is a MinA $\mathcal{M}$ for ϕ in $\mathcal{K}_G$. Furthermore, since G is minimal, for every $g \in G$ the axiom $a \to g$ is in $\mathcal{M}$. In particular, $\phi \in \mathcal{M}$.

($\Leftarrow$) Assume that there is such a MinA $\mathcal{M}$. It contains the axiom $\bigwedge_{m \in M} m \to b$, and also contains axioms from $\mathcal{C}$ such that every $m \in M$ occurs on the rhs of at least one axiom. Additionally $\mathcal{M}$ contains axioms of the form $a \to h$ such that $\mathcal{M} \models a \to \bigwedge_{m \in M} m$. Then the set $G := \{h \mid a \to h \in \mathcal{M}\}$ satisfies $G \cup \mathcal{C} \models M$. Moreover $p \in G$ since $a \to p \in \mathcal{M}$, and G is minimal since $\mathcal{M}$ is minimal. □

Concluding Remarks and Future Work

We have analyzed the complexity of axiom pinpointing and many related problems in the propositional Horn fragment and in the DL $\mathcal{EL}$. Since $\mathcal{EL}$ allows only for the constructors $\sqcap$ and $\exists$, our hardness results extend to any DL that includes these constructors. In some cases, hardness follows even without the existential restrictions. Table 2 summarizes our results, where TH stands for TRANS-HYP, TE stands for transversal enumeration, '-h' stands for hard, and '-c' stands for complete.

As future work we are going to work on determining the exact complexity of ALL-MINAS problem for dual-Horn KBs. That is, check whether it is equivalent to the TRANS-HYP problem, or strictly harder. We are also going to investigate the complexity of ALL-MINAS for more expressive DLs to see whether it remains in the same complexity class as reasoning. Moreover, we are going to look at the work in the SAT community on unsatisfiable core and other related problems, and investigate whether the methods developed there can be used for the problems we have discussed here. A different branch of future research is to look at the complexity of pinpointing in the DL-Lite family of DLs, where reasoning is also tractable like in the $\mathcal{EL}$ family of DLs.

Acknowledgements We are thankful to the anonymous reviewers for pointing out the connection to related work in the SAT community.

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