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First Order-Rewritability and Containment of Conjunctive Queries in Horn Description Logics

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Abstract

We study FO-rewritability of conjunctive queries in the presence of ontologies formulated in a description logic between $\mathcal{EL}$ and Horn-$\mathcal{SHIF}$, along with related query containment problems. Apart from providing characterizations, we establish complexity results ranging from $\text{ExpTime}$ via $\text{NExpTime}$ to $2\text{ExpTime}$, pointing out several interesting effects. In particular, FO-rewriting is more complex for conjunctive queries than for atomic queries when inverse roles are present, but not otherwise.

1 Introduction

When ontologies are used to enrich incomplete and heterogeneous data with a semantics and with background knowledge [Calvanese et al., 2009; Kontchakov et al., 2013; Bienvenu and Ortiz, 2015], efficient query answering is a primary concern. Since classical database systems are unaware of ontologies and implementing new ontology-aware systems that can compete with these would be a huge effort, a main approach used today is query rewriting: the user query $q$ and the ontology $\mathcal{O}$ are combined into a new query $\mathcal{Q}_0$ that produces the same answers as $q$ under $\mathcal{O}$ (over all inputs) and can be handed over to a database system for execution. Popular target languages for the query $\mathcal{Q}_0$ include SQL and Datalog. In this paper, we concentrate on ontologies formulated in description logics (DLs) and on rewriting into SQL, which we equate with first-order logic (FO).

FO-rewritability in the context of query answering under DL ontologies was first studied in [Calvanese et al., 2007]. Since FO-rewritings are not guaranteed to exist when ontologies are formulated in traditional DLs, the authors introduce the DL-Lite family of DLs specifically for the purpose of ontology-aware query answering using SQL database systems; in fact, the expressive power of DL-Lite is seriously restricted, in this way enabling existence guarantees for FO-rewritings. While DL-Lite is a successful family of DLs, there are many applications that require DLs with greater expressive power. The potential non-existence of FO-rewritings in this case is not necessarily a problem in practical applications. In fact, ontologies emerging from such applications typically use the available expressive means in a harmless way in the sense that efficient reasoning is often possible despite high worst-case complexity.

One might thus hope that, in practice, FO-rewritings can often be constructed also beyond DL-Lite.

This hope was confirmed in [Bienvenu et al., 2013; Hansen et al., 2015], which consider the case where ontologies are formulated in a DL of the $\mathcal{EL}$ family [Baader et al., 2005] and queries are atomic queries (AQs) of the form $A(x)$. To describe the obtained results in more detail, let an ontology-mediated query (OMQ) be a triple $(\mathcal{T}, \mathcal{Q}, q)$ with $\mathcal{T}$ a description logic TBox (representing an ontology), $\Sigma$ an ABox signature (the set of concept and role names that can occur in the data), and $q$ an actual query. Note that $\mathcal{T}$ and $q$ might use symbols that do not occur in $\Sigma$; in this way, the TBox enriches the vocabulary available for formulating $q$. We use $(\mathcal{L}, \mathcal{Q})$ to denote the OMQ language that consists of all OMQs where $\mathcal{T}$ is formulated in the description logic $\mathcal{L}$ and $q$ in the query language $\mathcal{Q}$. In [Bienvenu et al., 2013], FO-rewritability is characterized in terms of the existence of certain tree-shaped ABoxes, covering a range of OMQ languages between $(\mathcal{EL}, \mathcal{AQ})$ and (Horn-$\mathcal{SHIF}, \mathcal{AQ})$. On the one hand, this characterization is used to clarify the complexity of deciding whether a given OMQ is FO-rewritable, which turns out to be $\text{ExpTime}$-complete. On the other hand, it provides the foundations for developing practically efficient and complete algorithms for computing FO-rewritings. The latter was explored further in [Hansen et al., 2015], where a novel type of algorithm for computing FO-rewritings of OMQs from $(\mathcal{EL}, \mathcal{AQ})$ is introduced, crucially relying on the previous results from [Bienvenu et al., 2013]. Its evaluation shows excellent performance and confirms the hope that, in practice, FO-rewritings almost always exist. In fact, rewriting fails in only 285 out of 10989 test cases.

A limitation of the discussed results is that they concern only AQs while in many applications, the more expressive conjunctive queries (CQs) are required. The aim of the current paper is thus to study FO-rewritability of OMQ languages based on CQs, considering ontology languages between $\mathcal{EL}$ and Horn-$\mathcal{SHIF}$. In particular, we provide characterizations of FO-rewritability in the required OMQ languages that are inspired by those in [Bienvenu et al., 2013] (replacing tree-shaped ABoxes with a more general form of ABox), and we analyze the complexity of FO-rewritability using an automata-based approach. While practically efficient algorithms are out of the scope of this article, we believe that our work also lays important ground for the subsequent development of such...
algorithms. Our approach actually does allow the construction of rewritings, but it is not tailored towards doing that in a practically efficient way. It turns out that the studied FO-rewritability problems are closely related to OMQL containment problems as considered in [Bienvenu et al., 2012; Bourhis and Lutz, 2016]. In fact, being able to decide OMQL containment allows us to concentrate on connected CQs when deciding FO-rewritability, which simplifies technicalities considerably. For this reason, we also study characterizations and the complexity of query containment in the OMQL languages considered.

Our main complexity results are that FO-rewritability and containment are ExpTime-complete for OMQL languages between (EL, AQ) and (ELH(F_n), CQ) and 2ExpTime-complete for OMQL languages between (ELI, CQ) and (Horn-SHIF, CQ). The lower bound for containment applies already when both OMQLs share the same TBox. Replacing AQs with CQs thus results in an increase of complexity by one exponential in the presence of inverse roles (indicated by I), but not otherwise. Note that the effect that inverse roles can increase the complexity of querying-related problems was known from expressive DLs of the ALC family [Lutz, 2008], but it has not previously been observed for Horn-DLs such as ELI and Horn-SHIF. While 2ExpTime might appear to be very high complexity, we are fortunately also able to show that the runtime is double exponential only in the size of the actual queries (which tends to be very small) while it is only single exponential in the size of the ontologies. We also show that the complexity drops to NExpTime when we restrict our attention to rooted CQs, that is, CQs which contain at least one answer variable and are connected. Practically relevant queries are typically of this kind.

A slight modification of our lower bounds yields new lower bounds for monadic Datalog containment. In fact, we close an open problem from [Chaudhuri and Vardi, 1994] by showing that containment of a monadic Datalog program in a rooted CQ is coExpTime-complete. We also improve the 2ExpTime lower bound for containment of a monadic Datalog program in a CQ from [Benedikt et al., 2012] by showing that it already applies when the arity of EDB relations is bounded by two, rule bodies are tree-shaped, and there are no constants (which in this case correspond to nominals); the existing construction cannot achieve the latter two conditions simultaneously.

Full proofs are provided at http://www.informatik.unibremen.de/tkdi/research/papers.html.

Related work. Pragmatic approaches to OMQ rewriting beyond DL-Lite often consider Datalog as a target language [Rosati, 2007; Pérez-Urbina et al., 2010; Eiter et al., 2012; Kaminski et al., 2014; Trivela et al., 2015]. These approaches might produce a non-recursive (thus FO) rewriting if it exists, but there are no guarantees. FO-rewritability of OMQL is based on expressive DLs is considered in [Bienvenu et al., 2014], and based on existential rules in [Baget et al., 2011]. A problem related to ours is whether all queries are FO-rewritable when combined with a given TBox [Lutz and Wolter, 2012; Civili and Rosati, 2015]. There are several related works in the area of Datalog; recall that a Datalog program is bounded if and only if it is FO-rewritable [Ajtai and Gurevich, 1994]. For monadic Datalog programs, boundedness is known to be decidable [Cosmadakis et al., 1988] and 2ExpTime-complete [Benedikt et al., 2015]; containment is also 2ExpTime-complete [Cosmadakis et al., 1988; Benedikt et al., 2012]. OMQs from (Horn-SHIF, CQ) can be translated to monadic Datalog with an exponential blowup, functional roles (indicated by F) are not expressible.

2 Preliminaries and Basic Observations

Let \( N_c \) and \( N_b \) be disjoint and countably infinite sets of concept and role names. A role \( r \) is a role name \( r \) with an inverse role \( r^- \), with \( r \) a role name. A Horn-SHIF concept inclusion (CI) is of the form \( L \sqsubseteq R \), where \( L \) and \( R \) are concepts defined by the syntax rules

\[
R, R' ::= \top | \bot | A | \neg A | R \sqcap R' | L \sqcup R | \exists r . R | \forall r . R
\]

with \( A \) ranging over concept names and \( r \) over roles. In DLs, ontologies are formalized as TBoxes. A Horn-SHIF TBox \( T \) is a finite set of Horn-SHIF CIs, functionality assertions \( \text{func}(r) \), transitivity assertions \( \text{trans}(r) \), and role inclusions \( (RIs) r \sqsubseteq s \), with \( r \) and \( s \) roles. It is standard to assume that functional roles are not transitive and neither are transitive roles included in them (directly or indirectly). We make the slightly stronger assumption that functional roles do not occur on the right-hand side of role inclusions at all. This assumption seems natural from a modeling perspective and mainly serves the purpose of simplifying constructions; all our results can be extended to the milder standard assumption. An ELIHF TBox is a Horn-SHIF TBox that contains neither transitivity assertions nor disjunctions in CIs, an ELLI TBox is an ELIHF TBox that contains neither functionality assertions nor RIs, and an ELIFHF TBox is an ELIHF TBox that does not contain inverse roles.

An ABox is a finite set of concept assertions \( A(a) \) and role assertions \( r(a, b) \) where \( A \) is a concept name, \( r \) a role name, and \( a, b \) individual names from a countably infinite set \( N_i \). We sometimes write \( r^-(a, b) \) instead of \( r(b, a) \) and use \( \text{Ind}(A) \) to denote the set of all individual names used in \( A \). A signature is a set of concept and role names. We will often assume that the ABox is formulated in a prescribed signature, which we then call an ABox signature. An ABox that only uses concept and role names from a signature \( \Sigma \) is called a \( \Sigma \)-ABox.

The semantics of DLs is given in terms of interpretations \( I = (\Delta_I, -) \), where \( \Delta_I \) is a non-empty set (the domain) and \(-\) is the interpretation function, assigning to each \( A \in N_c \) a set \( A^I \subseteq \Delta_I \) and to each \( r \in N_b \) a relation \( r^I \subseteq \Delta_I \times \Delta_I \). The interpretation \( C^I \subseteq \Delta_I \times \Delta_I \) of a concept C in I is defined as usual, see [Baader et al., 2003]. An interpretation \( I \) satisfies a CI \( C \subseteq D \) if \( C^I \subseteq D^I \), a functionality assertion \( \text{func}(r) \) if \( r^I \) is a partial function, a transitivity assertion \( \text{trans}(r) \) if \( r^I \) is transitive, an RI \( r \sqsubseteq s \) if \( r^I \subseteq s^I \), a concept assertion \( A(a) \) if \( a \in A^I \), and a role assertion \( r(a, b) \) if \( (a, b) \in r^I \). We say that \( I \) is a model of a TBox or an ABox if it satisfies all inclusions and assertions in it. An ABox \( A \) is consistent with a TBox \( T \) if \( A \) and \( T \) have a common model. If \( A \) is a CI, RI, or functionality assertion, we write \( T \models A \) if all models of \( T \) satisfy \( A \).

A conjunctive query (CQ) takes the form \( q = \exists x. \varphi(x, y) \) with \( x, y \) tuples of variables and \( \varphi \) a conjunction of atoms of
the form \( A(x) \) and \( r(x,y) \) that uses only variables from \( x \cup y \). The variables in \( y \) are called answer variables, the arity of \( q \) is the length of \( y \), and \( q \) is Boolean if it has arity zero. An atomic query (AQ) is a conjunctive query of the form \( A(x) \). A union of conjunctive queries (UCQ) is a disjunction of CQs that share the same answer variables. Ontology-mediated queries (OMQs) and the notation \((\mathcal{L}, \mathcal{Q})\) for OMQ languages were already defined in the introduction. We generally assume that if a role name \( r \) occurs in \( q \) and \( T \models s \sqsubseteq r \), then \( \text{trans}(s) \notin T \). This is common since allowing transitive roles in the query poses serious additional complications, which are outside the scope of this paper; see e.g. [Bienvenu et al., 2010; Gottlob et al., 2013].

Let \( Q = (T, \Sigma, q) \) be an OMQ, \( q \) of arity \( n \), \( A \) a \( \Sigma \)-ABox and \( a \in \text{Ind}(A)^n \). We write \( A \models Q(a) \) if \( T \models q(a) \) for all models \( T \) of \( T \) and \( A \). In this case, \( a \) is a certain answer to \( Q \) on \( A \). We use \( \text{cert}(Q, A) \) to denote the set of all certain answers to \( Q \) on \( A \).

A first-order query (FOQ) is a first-order formula \( \varphi \) constructed from atoms \( A(x), r(x,y) \), and \( x = y \); here, concept names are viewed as unary predicates, role names as binary predicates, and predicates of other arity, function symbols, and constant symbols are not permitted. We write \( \varphi(x) \) to indicate that the free variables of \( \varphi \) are among \( x \) and call \( x \) the answer variables of \( \varphi \). The number of answer variables is the arity of \( \varphi \) and \( \varphi \) is Boolean if it has arity zero. We use \( \text{ans}(T, \varphi) \) to denote the set of answers to the FOQ \( \varphi \) on the interpretation \( T \); that is, if \( \varphi \) is \( n \)-ary, then \( \text{ans}(T, \varphi) \) contains all tuples \( d \in (\Delta^2)^m \) with \( T \models \varphi(d) \). To bridge the gap between certain answers and answers to FOQs, we sometime view an ABox \( A \) as an interpretation \( I_A \), defined in the obvious way.

For any syntactic object \( O \) (such as a TBox, a query, an OMQ), we use \( |O| \) to denote the size of \( O \), that is, the number of symbols needed to write it (concept and role names counted as a single symbol).

**Definition 1** (FO-rewriting). An FOQ \( \varphi \) is an FO-rewriting of an OMQ \( Q = (T, \Sigma, q) \) if \( \text{cert}(Q, A) = \text{ans}(I_A, \varphi) \) for all \( \Sigma \)-ABoxes \( A \) that are consistent with \( T \). If there is such a \( \varphi \), then \( Q \) is FO-rewritable.

**Example 2.** (1) Let \( Q_0 = (T_0, \Sigma_0, q_0(x,y)) \), where \( T_0 = \{ \exists r. A \sqsubseteq A, B \sqsubseteq \forall r.A \} \), \( \Sigma_0 = \{ r, A, B \} \) and \( q_0(x,y) = B(x) \wedge r(x,y) \wedge A(y) \). Then \( \varphi_0(x,y) = B(x) \wedge r(x,y) \) is an FO-rewriting of \( Q_0 \).

We will see in Example 10 that the query \( Q_A \) obtained from \( Q_0 \) by replacing \( q_0(x,y) \) with the AQ \( A(x) \) is not FO-rewritable (due to the unbounded propagation of \( A \) via \( r \)-edges by \( T_0 \)). Thus, an FO-rewritable OMQ can give raise to AQ subqueries with \( \varphi \).

(2) Let \( Q_1 = (T_1, \Sigma_1, q_1(x)) \), where \( T_1 = \{ \exists r. \exists A \sqsubseteq \exists r.A \} \), \( \Sigma_1 = \{ r, A \} \), and \( q_1(x) = \exists y(r(x,y) \wedge A(y)) \). Then \( Q_1 \) is not FO-rewritable (see again Example 10), but all AQ subqueries of \( Q_1 \) gives raise to are FO-rewritable.

The main reasoning problem studied in this paper is to decide whether a given OMQ \( Q = (T, \Sigma, q) \) is FO-rewritable. We assume without loss of generality that every symbol in \( \Sigma \) occurs in \( T \) or in \( q \). We obtain different versions of this problem by varying the OMQ language used. Note that we have defined FO-rewritability relative to ABoxes that are consistent with the TBox. It is thus important for the user to know whether that is the case. Therefore, we also consider FO-rewritability of ABox inconsistency. More precisely, we say that ABox inconsistency is FO-rewritable relative to a TBox \( T \) and ABox signature \( \Sigma \) if there is a Boolean FOQ \( \varphi \) such that for every \( \Sigma \)-ABox \( A \), \( A \) is inconsistent with \( T \) iff \( I_A \models \varphi(). \)

Apart from FO-rewritability questions, we will also study OMQ containment. Let \( Q_1 = (T_1, \Sigma_1, q_1) \) be two OMQs over the same ABox signature. We say that \( Q_1 \) is contained in \( Q_2 \), in symbols \( Q_1 \subseteq Q_2 \), if \( \text{cert}(Q_1, A) \subseteq \text{cert}(Q_2, A) \) holds for all \( \Sigma \)-ABoxes \( A \) that are consistent with \( T_1 \) and \( T_2 \).

We now make two basic observations that we use in an essential way in the remaining paper. We first observe that it suffices to concentrate on \( \mathcal{ELIF} \) TBoxes \( T \) and ABox signature \( \Sigma \), one can construct in polynomial time an \( \mathcal{ELIF} \) TBox \( T \) in normal form such that for every \( \Sigma \)-ABox \( A \).

1. \( A \) is consistent with \( T_1 \) if \( A \) is consistent with \( T_2 \);
2. if \( A \) is consistent with \( T_1 \), then for any CQ \( q \) that does not use symbols from \( \text{sig}(T_2) \setminus \text{sig}(T_1) \), we have \( \text{cert}(Q_1, A) = \text{cert}(Q_2, A) \) where \( Q_1 = (T_1, \Sigma, q) \).

Theorem 3 yields polyme time reductions of FO-rewritability in \( \mathcal{ELIF} \) \( Q \) to FO-rewritability in \( \mathcal{ELIF} \) \( Q \) for any query language \( Q \), and likewise for OMQ containment and FO-rewritability of ABox inconsistency. It also tells us that, when working with \( \mathcal{ELIF} \) TBoxes, we can assume normal form. Note that transitioning from \( \mathcal{ELIF} \) \( Q \) to \( \mathcal{ELIF} \) \( Q \) is not as easy as the case with inverse roles since universal restrictions on the right-hand side of concept inclusions cannot easily be eliminated: for this reason, we do not consider \( \mathcal{ELIF} \) \( Q \). From now on, we work with TBoxes formulated in \( \mathcal{ELIF} \) or \( \mathcal{ELIF} \) and assume without further notice that they are in normal form.

Our second observation is that, when deciding FO-rewritability, we can restrict our attention to connected queries provided that we have a way of deciding containment (for potentially disconnected queries). We use conCQ to denote the class of all connected CQs.

**Theorem 4.** Let \( L \in \{ \mathcal{ELIF}, \mathcal{ELIF} \} \). Then FO-rewritability in \( (\mathcal{L}, Q) \) can be solved in polynomial time when there is access to oracles for containment in \( (\mathcal{L}, Q) \) and for FO-rewritability in \( (\mathcal{L}, \text{conCQ}) \).

To prove Theorem 4, we observe that FO-rewritability of an OMQ \( Q = (T, \Sigma, q) \) is equivalent to FO-rewritability of all OMQs \( Q = (T, \Sigma, q_e) \) with \( q_e \) a maximal connected component of \( q \), excluding certain redundant such components (which can be identified using containment). Backed by Theorem 4, we generally assume connected queries when studying FO-rewritability, which allows to avoid unpleasant technical complications and is a main reason for studying FO-rewritability and containment in the same paper.
3 Main Results
In this section, we summarize the main results established in this paper. We start with the following theorem.

Theorem 5. FO-rewritability and containment are

1. 2EXPTime-complete for any OMQ language between \( \langle \Sigma T, CQ \rangle \) and \((\text{Horn-SHIFT}, CQ)\), and
2. EXPTime-complete for any OMQ language between \( \langle \Sigma A, CQ \rangle \) and \((\text{Horn-SHIFT}, CQ)\).

Moreover, given an OMQ from \((\text{Horn-SHIFT}, CQ)\) that is FO-rewritable, one can effectively construct a UCQ-rewriting.

Like the subsequent results, Theorem 5 illustrates the strong relationship between FO-rewritability and containment. Note that inverse roles increase the complexity of both reasoning tasks. We stress that this increase takes place only when the actual queries are conjunctive queries, since FO-rewritability for OMQ languages with inverse roles and atomic queries is in EXPTime [Bienvenu et al., 2013].

The 2EXPTime-completeness result stated in Point 1 of Theorem 5 might look discouraging. However, the situation is not quite as bad as it seems. To show this, we state the upper bound underlying Point 1 of Theorem 5 a bit more carefully.

Theorem 6. Given OMQs \( Q_i = (T_i, \Sigma_i, q_i), i \in \{1, 2\} \), from \((\text{Horn-SHIFT}, CQ)\), it can be decided

1. in time \( 2^{2^{O(1+\log(|T_1|))}} \) whether \( Q_1 \) is FO-rewritable and
2. in time \( 2^{2^{O(1+\log(|T_1|)+|T_2|)}} \) whether \( Q_1 \subseteq Q_2 \),

for some polynomial \( p \).

No note that the runtime is double exponential only in the size of the actual queries \( q_1 \) and \( q_2 \), while it is only single exponential in the size of the TBoxes \( T_1 \) and \( T_2 \). This is good news since the size of \( q_1 \) and \( q_2 \) is typically very small compared to the sizes of \( T_1 \) and \( T_2 \). For this reason, it can even be reasonable to assume that the sizes of \( q_1 \) and \( q_2 \) are constant, in the same way in which the size of the query is assumed to be constant in data complexity. Note that, under this assumption, Theorem 6 yields EXPTime upper bounds.

One other way to relativize the seemingly very high complexity stated in Point 1 of Theorem 5 is to observe that the lower bound proofs require the actual query to be Boolean or disconnected. In practical applications, though, typical queries are connected and have at least one answer variable. We call such CQs rooted and use rCQ to denote the class of all rooted CQs. Our last main result states that, when we restrict our attention to rooted CQs, then the complexity drops to coNEXPTime.

Theorem 7. FO-rewritability and containment are conEXPTime-complete in any OMQ language between \( \langle \Sigma T, rCQ \rangle \) and \((\text{Horn-SHIFT}, rCQ)\).

4 Semantic Characterization
The upper bounds stated in Theorems 5 and 6 are established in two steps. We first give characterizations of FO-rewritability in terms of the existence of certain (almost) tree-shaped ABoxes, and then utilize this characterization to design decision procedures based on alternating tree automata. The semantic characterizations are of independent interest.

An ABox \( A \) is tree-shaped if the undirected graph with nodes \( \text{Ind}(A) \) and edges \( \{\{a, b\} \mid r(a, b) \in A\} \) is acyclic and connected and \( r(a, b) \in A \) for all \( s \neq r \) and \( s(b, a) \notin A \) for all role names \( s \). For tree-shaped ABoxes \( A \), we often distinguish an individual used as the root, denoted with \( \rho_A \). \( A \) is ditree-shaped if the directed graph with nodes \( \text{Ind}(A) \) and edges \( \{(a, b) \mid r(a, b) \in A\} \) is a tree and \( r(a, b) \in A \) implies \( i \) and \( ii \). The (unique) root of a ditree-shaped ABox \( A \) is also denoted with \( \rho_A \). An ABox \( A \) is a pseudo tree if it is the union of ABoxes \( A_0, \ldots, A_k \) that satisfy the following conditions:

1. \( A_1, \ldots, A_k \) are tree-shaped;
2. \( k \leq |\text{Ind}(A_0)| \);
3. \( A_i \cap A_0 = \{\rho_A\} \) and \( \text{Ind}(A_i) \cap \text{Ind}(A_j) = \emptyset \), for \( 1 \leq i < j \leq k \).

We call \( A_0 \) the core of \( A \) and \( A_1, \ldots, A_k \) the trees of \( A \). The width of \( A \) is \( |\text{Ind}(A_0)| \), its depth is the depth of the deepest \( T \), and its outdegree is the maximum outdegree of the ABoxes \( A_1, \ldots, A_k \). For a pseudo tree ABox \( A \) and \( \ell \geq 0 \), we write \( A|_{\leq \ell} \) to denote the restriction of \( A \) to the individuals whose minimal distance from a core individual is at most \( \ell \), and analogously for \( A|_\leq \). A pseudo ditree ABox is defined analogously to a pseudo tree ABox, except that \( A_1, \ldots, A_k \) must be ditree-shaped.

When studying FO-rewritability and containment, we can restrict our attention to pseudo tree ABoxes, and even to pseudo ditree ABoxes when the TBox does not contain inverse roles. The following statement makes this precise for the case of containment. Its proof uses unraveling and compactness.

Proposition 8. Let \( Q_i = (T_i, \Sigma, q_i), i \in \{1, 2\} \), be OMQs from \((\text{Horn-SHIFT}, CQ)\). Then \( Q_1 \subseteq Q_2 \) iff there is a pseudo tree \( \Sigma-ABox A \) of outdegree at most \( |T_1| \) and width at most \( |q_1| \) that is consistent with both \( T_1 \) and \( T_2 \) and a tuple \( a \) from the core of \( A \) such that \( A \models Q_1(a) \) and \( A \not\models Q_2(a) \).

If \( Q_1, Q_2 \) are from \((\text{Horn-SHIFT}, CQ)\), then we can find a pseudo ditree ABox with these properties.

We now establish a first version of the announced characterizations of FO-rewritability. Like Proposition 8, they are based on pseudo tree ABoxes.

Theorem 9. Let \( Q = (T, \Sigma, q) \) be an OMQ from \((\text{Horn-SHIFT}, \Sigma-CQ)\).

If the arity of \( q \) is at least one, then the following conditions are equivalent:

1. \( Q \) is FO-rewritable;
2. there is a \( k \geq 0 \) such that for all pseudo tree \( \Sigma-ABoxes A \) that are consistent with \( T \) and of outdegree at most \( |T| \) and width at most \( |q| \): \( \text{if } A \models Q(a) \text{ with } a \text{ from the core of } A, \text{ then } A|_{\leq k} \models Q(a) \).

If \( q \) is Boolean, this equivalence holds with (2.) replaced by

2. there is a \( k \geq 0 \) such that for all pseudo tree \( \Sigma-ABoxes A \) that are consistent with \( T \) and of outdegree at most \( |T| \) and width at most \( |q| \): \( \text{if } A \models Q, \text{ then } A|_{\geq 0} \models Q \text{ or } A|_{\leq k} \models Q \).

If \( Q \) is from \((\text{Horn-SHIFT}, \Sigma-CQ)\), then the above equivalences hold also when pseudo tree \( \Sigma-ABoxes \) are replaced with pseudo ditree \( \Sigma-ABoxes \).
The proof of Proposition 8 gives a good intuition of why FO-rewritability can be characterized in terms of ABoxes that are pseudo trees. In fact, the proof of “$2 \Rightarrow 1$” of Theorem 9 is similar to the proof of Proposition 8. The proof of “$1 \Rightarrow 2$” uses locality arguments in the form of Ehrenfeucht-Fraïssé games. The following examples further illustrate Theorem 9.

**Example 10.** (1) Non FO-rewritability of the OMQs $Q_A$ and $Q_1$ from Example 2 is shown by refuting Condition 2 in Theorem 9: let $A_k = \{r(a_0, a_1), \ldots, r(a_k, a_{k+1}), A(a_{k+1})\}$, for all $k \geq 0$. Then $A_k \not\models Q(a_0)$ but $A_k \models Q(a_0)$ for $Q \in \{Q_A, Q_1\}$.

(2) Theorem 9 only holds for connected CQs: consider $Q_2 = (\Sigma_2, \Sigma_2, q_2)$, where $\Sigma_2$ is the empty TBox, $\Sigma_2 = \{A, B\}$, and $q_2 = \exists x \forall y (A(x) \land B(y))$. $Q_2$ is FO-rewritable ($q_2$ itself is a rewriting), but Condition 2’ does not hold: for $B_k = \{A(a_0), R(a_0, a_1), \ldots, R(a_k, a_{k+1}), B(a_{k+1})\}$ we have $B_k \models Q_2$ but $B_k \models Q_2$ and $B_k \not\models Q_2$.

(3) The modifiﬁcation 2’ of Condition 2 is needed to charac-
terize FO-rewritability of Boolean OMQs: obtain $Q_3$ from $Q_2$ by replacing $q_2$ with $\exists x B(x)$. Then $Q_3$ is FO-rewritable, but the ABoxes $B_k$ show that Condition 2 does not hold.

Theorem 9 does not immediately suggest a decision pro-
cedure for FO-rewritability since there is no bound on the depth of the pseudo tree ABoxes $A$ used. The next result establishes such a bound.

**Theorem 11.** Let $T$ be an $\mathcal{ELHI} F L_1$ TBox. Then Theorem 9 still holds with the following modifiﬁcations:

1. if $q$ is not Boolean or $T$ is an $\mathcal{ELHI} F L_1$ TBox, “there is a $k \geq 0$” is replaced with “for $k = |q| + 2^{|T|+|q|^2}$”;

2. if $q$ is Boolean, “there is a $k \geq 0$” is replaced with “for $k = |q| + 2^{|T|+2^{|q|}}$”.

The proof of Theorem 11 uses a pumping argument based on derivations of concept names in the pumped ABox by $T$. Due to the presence of inverse roles, this is not entirely triv-
ial and uses what we call transfer sequences, describing the derivation history at a point of an ABox. Together with the proof of Theorem 9, Theorem 11 gives rise to an algorithm that constructs actual rewritings when they exist.

## 5 Constructing Automata

We show that Proposition 8 and Theorem 11 give rise to automata-based decision procedures for containment and FO-rewritability that establish the upper bounds stated in Theo-
rems 5 and 6. By Theorem 4, it sufﬁces to consider connected queries in the case of FO-rewritability. We now observe that we can further restrict our attention to Boolean queries. We use BCQ (resp. conBCQ) to denote the class of all Boolean CQs (resp. connected Boolean CQs).

**Lemma 12.** Let $\mathcal{L} \in \{\mathcal{ELHI} F L_1, \mathcal{ELH} F L_1\}$. Then

1. FO-rewritability in $(\mathcal{L}, \text{conCQ})$ can be reduced in poly-
time to FO-rewritability in $(\mathcal{L}, \text{conBCQ})$;

2. Containment in $(\mathcal{L}, CQ)$ can be reduced in polytime to containment in $(\mathcal{L}, \text{BCQ})$.

The decision procedures rely on building automata that accept pseudo tree ABoxes which witness non-containment and non-FO-rewritability as stipulated by Proposition 8 and Theorem 11, respectively. We first have to encode pseudo tree ABoxes in a suitable way.

A tree is a non-empty (and potentially inﬁnite) set $T \subseteq \mathbb{N}^*$ closed under prefixes. We say that $T$ is $m$-ary if for every $x \in T$, the set $\{i \mid x \cdot i \in T\}$ is of cardinality at most $m$. For an alphabet $\Gamma$, a $\Gamma$-labeled tree is a pair $(T, L)$ with $T$ a tree and $L : T \rightarrow \Gamma$ a node labeling function. Let $Q = (T, \Sigma, q)$ be an OMQ from $(\mathcal{ELHI} F L_1, \text{conBCQ})$. We encode pseudo ABoxes of width at most $|q|$ and outdegree at most $|T|$ by $(|T| \cdot |q|)$-ary $\Sigma_2 \cup \Sigma_N$-labeled trees, where $\Sigma_2$ is an alphabet used for labeling root nodes and $\Sigma_N$ is for non-root nodes.

The alphabet $\Sigma_2$ consists of all $\Sigma$-ABoxes $A$ such that Ind$(A)$ only contains individual names from a ﬁxed set Indcore of size $|q|$ and $A$ satisﬁes all functionality statements in $T$. The alphabet $\Sigma_N$ consists of all subsets $\Theta \subseteq (\mathbb{N}_C \cap \Sigma) \cup \{r, r^-, r \in \mathbb{N}_R \cap \Sigma\} \cup \text{Indcore}$ that contain exactly one (potentially inverse) role and at most one element of Indcore. A $(|T| \cdot |q|)$-ary $\Sigma_2 \cup \Sigma_N$-labeled tree is proper if (i) the root node is labeled with a symbol from $\Sigma_2$, (ii) each child of the root is labeled with a symbol from $\Sigma_N$ that contains an element of Indcore, (iii) every other non-root node is labeled with a symbol from $\Sigma_N$ that contains no individual name, and (iv) every non-root node has at most $|q|$ successors and (v) for every $a \in \text{Indcore}$, the root node has at most $|q|$ successors whose label includes $a$.

A proper $\Sigma_2 \cup \Sigma_N$-labeled tree $(T, L)$ represents a pseudo tree ABox $A(T, L)$ whose individuals are those in the ABox $A$ that labels the root of $T$ plus all non-root nodes of $T$, and whose assertions are

$$\forall A \in L(x) \land A \subseteq L(x)$$

$$\forall \{r(b, x) \mid \{b, r\} \subseteq L(x)\} \cup \{r(b, x) \mid \{b, r^-\} \subseteq L(x)\}$$

$$\forall \{r(x, y) \mid r \in \{L(y), y \text{ is a child of } x, L(x) \subseteq \Sigma_N\}$$

$$\forall \{r(x, y) \mid r^- \in \{L(y), y \text{ is a child of } x, L(x) \subseteq \Sigma_N\}$$

As the automaton model, we use two-way alternating parity automata on ﬁnite trees (TWAPAs). As usual, $L(\mathfrak{A})$ denotes the tree language accepted by the TWAPA $\mathfrak{A}$. Our central observation is the following.

**Proposition 13.** For every OMQ $Q = (T, \Sigma, q)$ from $(\mathcal{ELHI} F L_1, \text{BCQ})$, there is a TWAPA

1. $\mathfrak{A}_Q$ that accepts a $(|T| \cdot |q|)$-ary $\Sigma_2 \cup \Sigma_N$-labeled tree $(T, L)$ if it is proper, $\mathfrak{A}(T, L) = \text{consistent with } T$, and $A(T, L) \models Q$;

2. $\mathfrak{A}_T$ that accepts a $(|T| \cdot |q|)$-ary $\Sigma_2 \cup \Sigma_N$-labeled tree $(T, L)$ if it is proper and $A(T, L) = \text{consistent with } T$.

$\mathfrak{A}_T$ has at most $p(|T|)$ states, $p$ a polynomial.

We can construct $\mathfrak{A}_Q$ and $\mathfrak{A}_T$ in time polynomial in their size.

The construction of the automata in Proposition 13 uses forest decompositions of the OMQ $q$ as known for example from [Lutz, 2008]. The difference in automata size between $\mathcal{ELHI} F L_1$ and $\mathcal{ELHI} F L_1$ is due to the different number of tree-shaped subqueries that can arise in these decompositions.
To decide $Q_1 \subseteq Q_2$ for OMQs $Q_i = (\mathcal{T}_i, \Sigma, q_i)$, $i \in \{1, 2\}$, from $(\mathcal{ELTHF}_\bot, \text{BCQ})$, by Proposition 8 it suffices to decide whether $L(\mathcal{A}_{Q_1}) \cap L(\mathcal{T}_2) \subseteq L(\mathcal{A}_{Q_2})$. Since this question can be polynomially reduced to a TWAPA emptiness check and the latter can be executed in time single exponential in the number of states, this yields the upper bounds for containment stated in Theorems 5 and 6.

To decide non-FO-rewritability of an OMQ $Q = (\mathcal{T}, \Sigma, q)$ from $(\mathcal{ELTHF}_\bot, \text{conBCQ})$, by Theorem 11 we need to decide whether there is a pseudo tree $\Sigma$-ABox $\mathcal{A}$ of outdegree at most $|\mathcal{T}|$ and width at most $|q|$ that is consistent with $\mathcal{T}$ and satisfies (i) $\mathcal{A} \models Q$, (ii) $|\mathcal{A}|_{>0} \not\models Q$, and (iii) $|\mathcal{A}|_{\leq k} \not\models Q$ where $k = |q| + 2^{\kappa(|\mathcal{T}|+2^n)^2}$. For consistency with $\mathcal{T}$ and for (i), we use the automaton $\mathcal{A}_Q$ from Proposition 13. To achieve (ii) and (iii), we amend the tree alphabet $\Sigma_e \cup \Sigma_a$ with additional labels that implement a counter which counts up to $k$ and annotate each node in the tree with its depth (up to $k$).

The results for containment apply already when both OMQs share the same TBox.

Theorem 16. Containment and FO-rewritability are
1. $\text{coNEXPTime-hard in } (\mathcal{EL}, r\text{CQ})$ and
2. $\text{2EXPTime-hard in } (\mathcal{EL}, \text{CQ})$.

The proof of Point 2 of Theorem 16 can be modified to yield new lower bounds for monadic Datalog containment. Recall that the rule body of a Datalog program is a CQ, Tree-shapenedness of a CQ $q$ is defined in the same way as for an ABox in Section 4, that is, $q$ viewed as an undirected graph must be a tree without multi-edges.
References


