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► **To cite this version:**

M. Bojanczyk. Data Monoids. Symposium on Theoretical Aspects of Computer Science (STACS2011), Mar 2011, Dortmund, Germany. pp.105-116. hal-00573630

HAL Id: hal-00573630

<https://hal.science/hal-00573630v1>

Submitted on 5 Mar 2011

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Data Monoids*

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Abstract

We develop an algebraic theory for languages of data words. We prove that, under certain conditions, a language of data words is definable in first-order logic if and only if its syntactic monoid is aperiodic.

1998 ACM Subject Classification F.4.3 Formal Languages

Keywords and phrases Monoid, Data Words, Nominal Set, First-Order Logic

Digital Object Identifier 10.4230/LIPIcs.STACS.2011.105

1 Introduction

This paper is an attempt to combine two fields.

The first field is the algebraic theory of regular languages. In this theory, a regular language is represented by its syntactic monoid, which is a finite monoid. It turns out that many important properties of the language are reflected in the structure of its syntactic monoid. One particularly beautiful result is the Schützenberger-McNaughton-Papert theorem, which describes the expressive power of first-order logic.

Let $L \subseteq A^*$ be a regular language. Then L is definable in first-order logic if and only if its syntactic monoid M_L is aperiodic.

For instance, the language “words where there exists a position with label a ” is defined by the first-order logic formula (this example does not even use the order on positions $<$, which is also allowed in general)

$$\exists x. a(x).$$

The syntactic monoid of this language is isomorphic to $\{0, 1\}$ with multiplication, where 0 corresponds to the words that satisfy the formula, and 1 to the words that do not. Clearly, this monoid does not contain any non-trivial group. There are many results similar to theorem above, each one providing a connection between seemingly unrelated concepts of logic and algebra, see e.g. the book [8].

The second field is the study of languages over infinite alphabets, commonly called languages of data words. Regular languages are usually considered for finite alphabets. Many current motivations, including XML and verification, require the use of infinite alphabets. In this paper, we use the following definition of data words. We fix for the rest of the paper a single infinite alphabet D , and study words and languages over D . A typical language, which we use as our running example, is “some two consecutive positions have the same letter”, i.e.

$$L_{dd} = \bigcup_{d \in D} D^* dd D^* = \{d_1 \cdots d_n \in D^* : d_i = d_{i+1} \text{ for some } i \in \{1, \dots, n-1\}\}.$$

* Author supported by ERC Starting Grant “Sosna”.



A number of automata models have been developed for such languages. The general theme is that there is a tradeoff between the following three properties: expressivity, good closure properties, and decidable (or even efficiently decidable) emptiness. The existing models strike different balances in this tradeoff, and it is not clear which balance, if any, should deserve the name of “regular language”. Also, logics have been developed to express properties of data words. For more information on automata and logics for data words, see the survey [7].

The motivation of this paper is to combine the two fields, and prove a theorem that is analogous to the Schützenberger-McNaughton-Papert characterization theorem, but talks about languages of data words. If we want an analogue, we need to choose the notion of regular language for data words, the definition of first-order logic for data words, and then find the property corresponding to aperiodicity.

For first-order logic over data words we use a notion that has been established in the literature. Formulas are evaluated in data words. The quantifiers quantify over positions of the data word, we allow two binary predicates: $x < y$ for order on positions, and $x \sim y$ for having the same data value. An example formula of a formula of first-order logic is the formula

$$\exists x \exists y \quad x < y \quad \wedge \quad x \sim y \quad \wedge \quad \neg(\exists z \ x < z \wedge z < y),$$

which defines the language L_{dd} in the running example.

As for the notion of regular language, we use the monoid approach. For languages of data words, we use the syntactic monoid, defined in the same way as it is for words over finite alphabets. That is, elements of the syntactic monoid are equivalence classes of the two-sided Myhill-Nerode congruence. When the alphabet is infinite, the syntactic monoid is infinite for every language of data words, except those that depend only on the length of the word. For instance, in the case of the running example language L_{dd} , two words $w, w' \in D^*$ are equivalent if and only if: either both belong to L_{dd} , or both have the same first and last letters. Since there are infinitely many letters, there are infinitely many equivalence classes.

Instead of studying finite monoids, we study something called *orbit finite* monoids. Intuitively speaking, two elements of a syntactic monoid are considered to be in the same orbit if there is a renaming of data values that maps one element to the other¹. For instance, in the running example L_{dd} , the elements of the syntactic monoid that correspond to the words $1 \cdot 7$ and $2 \cdot 3 \cdot 4 \cdot 8$ are not equal, but they are in the same orbit, because the renaming $i \mapsto i + 1$ maps $1 \cdot 7$ to $2 \cdot 8$, which corresponds to the same element of the syntactic monoid as $2 \cdot 3 \cdot 4 \cdot 8$. It is not difficult to see that the syntactic monoid of L_{dd} has four orbits: one for the empty word, one for words inside L_{dd} , one for words outside L_{dd} where the first and last letters are equal, and one for words outside L_{dd} where the first and last letters are different.

The contribution of this paper is a study of orbit finite data monoids. We develop the algebraic theory of orbit finite data monoids, and show that it resembles the theory of finite monoids. The main result of the paper is Theorem 11, which shows that the Schützenberger-McNaughton-Papert characterization also holds for languages of data words with orbit finite syntactic data monoids:

Let $L \subseteq D^*$ be a language whose syntactic data monoid M_L is orbit finite. Then L is definable in first-order logic if and only if M_L is aperiodic.

¹ This is related to Proposition 3 in [6]

Related work. The idea to present effective characterizations of logics on data words was proposed by Benedikt, Ley and Puppis. In [1], they show that definability in first-order logic is undecidable if the input is a nondeterministic register automaton. Also, they provide some decidable characterizations, including a characterization of first-order logic with local data comparisons within the class of languages recognized by deterministic register automata. (This result is incomparable the one in this paper.)

The key property of data monoids, namely that elements of the monoid are acted on by permutations of data values, appears in the nominal sets of Gabbay and Pitts [5]. The connections with the literature on nominal sets are still relatively unexplored.

There are two papers on algebra for languages over infinite alphabets. One approach is described by Bouyer, Petit and Thérien [3]. Their monoids are different from ours in the following ways: the definition of a monoid explicitly talks about registers, there is no syntactic monoid, monoids have undecidable emptiness (not mention questions such as aperiodicity or first-order definability). Our setting is closer to the approach of Francez and Kaminski from [4], but the latter talks about automata and not monoids, and does not study the connection with logics.

Acknowledgments. I would like to thank Clemens Ley for introducing me to the subject, Sławomir Lasota for suggesting the key idea of orbits, and Bartek Klin for pointing out the connection to nominal sets. I would also like to thank Michael Kaminski and the anonymous referees for their many helpful comments.

2 Myhill-Nerode equivalence

We start out with an investigation of the Myhill-Nerode syntactic congruence, in the case of data words.

We fix for the rest of this paper an infinite set of data values D . One can think of D as being the natural numbers, but all the structure of natural numbers (such as order) is not going to play a role. Data words are words in D^* . Inside D , we will distinguish a finite subset $C \subseteq D$ of constants. Although D is fixed, the constants will vary. The constants are introduced for two purposes. First, interpreting a finite alphabet as a set of constants, one recovers the traditional theory of regular languages in this setting. Second, constants are useful in the proofs. All definitions will use the set of constant as a parameter.

We use the term *data renaming (with constants C)* for a bijection on the alphabet that is the identity function on C .² To simplify proofs, we require a data renaming to be the identity on all but finitely many data values³. We write Γ_C for the set of all data renamings with constants C . This set has a group structure. We write π, τ, σ for data renamings. When we apply a data renaming π to a data value d , we skip the brackets and write πd instead of $\pi(d)$.

A data renaming π is naturally extended to a homomorphism $[\pi] : D^* \rightarrow D^*$ of data words. For a fixed set of constants C , we study only languages that are invariant under data renaming, i.e. languages such that $[\pi]L = L$ for any data renaming $\pi \in \Gamma_C$.⁴

Suppose that $L \subseteq D^*$ is a language that is invariant under data renamings. We define two equivalence relations on D^* .

² This is called C -preservation in Definition 3 of [4].

³ The same results hold when all bijections are allowed, including the Memory Theorem.

⁴ These languages are called co- C -invariant languages in Definition 8 of [4].

The first equivalence relation is the two-sided Myhill-Nerode equivalence relation:

$$w \equiv_L v \quad \text{if for all } u_1, u_2 \in D^*, u_1 w u_2 \in L \Leftrightarrow u_1 v u_2 \in L.$$

We also use the name *syntactic congruence* for this relation. If L uses only constants, i.e. $L \subseteq C^*$, then this equivalence relation has finitely many equivalence classes. However, for any language that uses non-constants in a nontrivial way, this equivalence relation will have infinitely many equivalence classes. The problem is that the relation really cares about specific data values.

This problem is fixed by the second equivalence relation:

$$w \simeq_L v \quad \text{if for some data renaming } \tau \in \Gamma_C, [\tau]w \equiv_L v.$$

As discussed in the introduction, in the running example L_{dd} this equivalence relation has four equivalence classes. On the other hand, some reasonable languages have infinitely many equivalence classes in \simeq_L . An example of such an L is “the first letter also appears one some other position”. This language is recognized by a left-to-right register automaton with three states. Nevertheless, an equivalence class of a word w under \equiv_L is determined by the first letter of w and the set of distinct letters in w . Consequently, any two words with a different number of distinct letters are not equivalent under \simeq_L . (This problem disappears when one uses the one-sided Myhill-Nerode equivalence.)

Despite the limitations, we will be studying languages that have finitely many equivalence classes for \simeq_L , because they yield a monoid structure.

3 Data monoids

Fix a finite set of constants $C \subseteq D$. We want to define something like a monoid for data languages. We want there to be a “free object”, which should describe D^* . We want each language to have its syntactic object, which should be finite in many interesting cases, and be somehow optimal for all other objects capturing the language. These goals are achieved by the notion of a data monoid, as defined below.

Definition of a data monoid. A *data monoid* (with constants C) is a monoid M , which is acted upon by data renamings (with constants C). That is, for every data renaming $\tau \in \Gamma_C$, there is an associated monoid automorphism $[\tau] : M \rightarrow M$. The mapping $\tau \mapsto [\tau]$ should be a group action, namely $[\tau \circ \sigma] = [\tau] \circ [\sigma]$, and the identity renaming should be associated to the identity automorphism. There are two key examples:

- The free data monoid. The monoid is D^* , while the action is defined by

$$[\tau](d_1 \cdots f_n) = \tau(d_1) \cdots \tau(d_n) \quad \text{for } d_1, \dots, d_n \in D.$$

- The syntactic data monoid of a language $L \subseteq D^*$, which we denote M_L . The monoid is the quotient of D^* under the Myhill-Nerode equivalence \equiv_L . This is a well defined monoid, since \equiv_L is a monoid congruence (unlike \simeq_L). The action is defined by

$$[\tau]([w]_{\equiv_L}) = [[\tau](w)]_{\equiv_L}.$$

In Lemma 1, we show that this action is well defined and does not depend on w .

In the sequel, we will drop the brackets $[\tau]$ and simply write τ , when the context indicates if τ is treated as function on data values, or on elements of the monoid. Note that for two renamings τ, σ and an element m of the monoid, the notation $\tau\sigma m$ is unambiguous, since we are dealing with an action.

Finite support axiom. We say that a set $X \subseteq D$ of data values *supports* an element m of a data monoid, if $\tau m = m$ holds for every data renaming in $\tau \in \Gamma_X$. We require an additional property from a data monoid, called the *finite support axiom*: for any element m of the data monoid, there is a finite support.

Both examples above satisfy the finite support axiom. The axiom is violated by some data monoids, even finite ones, as shown in the following example, due to Szymon Toruńczyk. Let \mathbb{Z}_2 be the two element group, with additive notation. Let $h : \Gamma_\emptyset \rightarrow \mathbb{Z}_2$ assign 0 to even permutations and 1 to odd permutations. The monoid M is also \mathbb{Z}_2 ; the action is defined by $\tau m = m + h(\tau)$. The finite support axiom fails, since odd permutations can involve any pair of data values.

Nevertheless, all data monoids in this paper are homomorphic images (see below) of the free data monoid, which guarantees the finite support axiom.

Congruences of data monoids. A congruence of a data monoid is an equivalence relation \simeq on the underlying monoid that is compatible with concatenation:

$$m_1 \simeq n_1 \text{ and } m_2 \simeq n_2 \quad \text{implies} \quad m_1 n_1 \simeq m_2 n_2 \quad (1)$$

and which is also compatible with every data renaming $\tau \in \Gamma_C$:

$$m \simeq n \quad \text{implies} \quad \tau m \simeq \tau n \quad (2)$$

Note that the notion of congruence depends implicitly on the set C of constants, since the set of constants indicates which functions τ are data renamings. It might be the case that a relation is a congruence for a set of constants C , but it is no longer a congruence for a smaller set of constants $E \subsetneq C$.

► **Lemma 1.** *For a data language $L \subseteq D^*$ that is closed under data renamings, the syntactic equivalence \equiv_L is a congruence of the free data monoid D^* .*

Proof. The syntactic equivalence satisfies (1), as it does for any language $L \subseteq D^*$, not only those closed under data renamings. We use the closure under data renamings to prove (2). Let w and w' be \equiv_L -equivalent data words. We need to show that for any data renaming τ , also τw and $\tau w'$ are \equiv_L -equivalent. In other words, we need to show that

$$u(\tau w)v \in L \quad \Leftrightarrow \quad u(\tau w')v \in L \quad \text{for any } u, v \in D^*. \quad (3)$$

Since τ^{-1} induces a bijection on data words, we have

$$u(\tau w)v \in L \quad \Leftrightarrow \quad \tau^{-1}(u(\tau w)v) \in \tau^{-1}L.$$

Because concatenating words commutes with data renamings,

$$\tau^{-1}(u(\tau w)v) = (\tau^{-1}u)w(\tau^{-1}v).$$

By the assumption on L being closed under data renamings, $L = \tau^{-1}L$. It follows that the left side of the equivalence (3) is equivalent to

$$(\tau^{-1}u)w(\tau^{-1}v) \in L.$$

Likewise, the right side of the equivalence (3) becomes

$$(\tau^{-1}u)w'(\tau^{-1}v) \in L.$$

Both sides are equivalent, because $w \equiv_L w'$. ◀

Homomorphisms of data monoids. Suppose that M and N are data monoids, with constants C . A *data monoid homomorphism* $h : M \rightarrow N$ is defined to be a monoid homomorphism of the underlying monoids that commutes with the action of data renamings

$$h(\tau s) = \tau(h(s)) \quad \text{for all } \tau \in \Gamma_C.$$

In the literature on nominal sets, a monoid homomorphism that satisfies the commuting condition above is called *equivariant*.

A congruence gives rise to a homomorphism, in the usual way. It follows that the function that maps a word w to its equivalence class under \equiv_L , is a data monoid homomorphism. This function is called the *syntactic morphism*. The target of the syntactic morphism is called the *syntactic data monoid* of L , and denoted M_L .

Note that data monoid homomorphisms preserve the finite support axiom. Therefore, all the syntactic monoids we will study, as images of the free data monoid, will satisfy the finite support axiom.

We say that a data monoid homomorphism $h : D^* \rightarrow M$ *recognizes* a data language L if membership $w \in L$ is uniquely determined by the image $h(w) \in M$. In other words, $L = h^{-1}(h(L))$. Clearly, the syntactic morphism of L recognizes L . The following lemma shows that the syntactic morphism is the “best” morphism recognizing L .

► **Lemma 2.** *Consider a data language $L \subseteq D^*$. Any surjective data monoid homomorphism that recognizes L can be uniquely extended to the syntactic data monoid homomorphism.*

Proof. Let $\alpha : D^* \rightarrow M$ be a surjective data monoid homomorphism that recognizes L .

We claim that $\alpha(w) = \alpha(w')$ implies $w \equiv_L w'$. If the contrary were true, we would have two words w, w' with the same image under α , and some words v, u such that exactly one of the two words $vwu, vw'u$ would belong to L . This would contradict the assumption on α recognizing L , since $vwu, vw'u$ would have the same image under α .

From the claim, it follows that for every element $m \in M$, all words in $\alpha^{-1}(m)$ are mapped by the syntactic monoid homomorphism to the same equivalence class, call it W_m . Therefore, the function $m \mapsto W_m$ extends α to the syntactic data monoid homomorphism. ◀

The following lemma justifies the use of the name “free data monoid”.

► **Lemma 3.** *Let M be a data monoid, and let $h : D \rightarrow M$ be a function that commutes with data renamings (i.e. an equivariant function). This function can be uniquely extended to a data monoid homomorphism $[h] : D^* \rightarrow M$.*

Orbits. A syntactic data monoid usually has an infinite carrier. This is not a problem, because what really interests us in a data monoid is not the elements, but their orbits. Formally, the *orbit* of an element m of a data monoid M is the set

$$\Gamma_C \cdot m = \{\tau m : \tau \in \Gamma_C\}.$$

Note how the set of constants influences the notion of orbit.

Consider the syntactic data monoid and the syntactic data monoid homomorphism of a data language $L \subseteq D^*$. As elements of the monoid are to equivalence classes of \equiv_L , orbits are to equivalence classes of \simeq_L . More formally, for two words $w, v \in D^*$, the orbits of $h_L(w)$ and $h_L(v)$ are equal if and only if $w \simeq_L v$. Suppose that a language $L \subseteq D^*$ is closed under data renamings. If this language is recognized by a data monoid homomorphism h , then the image $h(L)$ is necessarily a union of orbits.

A data monoid is called *orbit finite* if it has finitely many distinct orbits. We are interested in languages whose syntactic data monoids are orbit finite. As far as this paper is concerned, these are the “regular” languages of data words.

This completes the definition of data monoids and their basic properties.

4 Memory

As usual, the definition of data monoids attempts to use the least amount of primitive concepts. We now show how other natural concepts can be derived from these primitives.

Consider the running example language L_{dd} . Let M_{dd} be the syntactic monoid of this language. Formally speaking, elements of M_{dd} are equivalence classes of data words. The equivalence class of a data word $d_1 \dots d_n \notin L_{dd}$ can be identified with the ordered first/last letter pair (d_1, d_n) . Intuitively speaking, the data values d_1 and d_n are “important” for this equivalence class (monoid element), while the other data values d_2, \dots, d_{n-1} are not “important”. Below we present a definition, called *memory*, which automatically extracts the “important” data values from an element of a data monoid, and uses only the primitive concepts of a data monoid. This definition is inspired by the notion of memory from [2].

Let m be an element of a data monoid. We write stab_m for the subgroup of data renamings τ that satisfy $\tau m = m$. In other words, $\text{stab}_m = \Gamma_{C \cup \{m\}}$. In the running example, when m is the equivalence class of a word $d_1 \dots d_n \notin L_{dd}$, then stab_m is the set of data renamings that have both d_1 and d_n as fixpoints. (When m is the equivalence class of a word $w \in L_{dd}$, or m is the identity, then stab_m contains all data renamings.) For each data value $d \in D$, we define its *data orbit with respect to m* as the following set of data values

$$\{\tau d : \tau \in \text{stab}_m\} = \text{stab}_m(d).$$

In the running example, when m is the equivalence class of $d_1 \dots d_n \notin L_{dd}$, the data orbit of d_1 with respect to m is $\{d_1\}$, for d_n the data orbit is $\{d_n\}$, and for any other data value it is $D - \{d_1, d_n\}$.

The following theorem gives a more precise interpretation of the finite support axiom. It says that an element does not care about what happens when data values in its unique infinite data orbit are swapped around.

► **Theorem 4 (Memory Theorem).** *Every element m of an orbit finite data monoid has finitely many data orbits, of which exactly one is infinite. Furthermore, if $\text{mem}(m)$ is defined as the union of the finite data orbits, then stab_m contains all data renamings that are the identity on $\text{mem}(m)$, i.e. $\Gamma_{\text{mem}(m)} \subseteq \text{stab}_m$.*

Note that $\text{mem}(m)$ includes all constants, since they have one element data orbits.

The result above only refers to the set structure of a data monoid, i.e. the group action on its elements. It does not refer to the monoid structure, i.e. the multiplication operation and the neutral element. Therefore, it is a result about nominal sets, as in [5]. The Memory Theorem is a corollary of Proposition 3.4 from [5], which says that an element m of a nominal set M is supported by the intersection of all sets of data values that support it.

► **Corollary 5.** *For m, n in a data monoid, $\text{mem}(mn) \subseteq \text{mem}(m) \cup \text{mem}(n)$.*

Proof. Let d, e be elements outside $\text{mem}(m) \cup \text{mem}(n)$. Let τ be the data renaming that swaps d, e . By the Memory Theorem, $\tau \in \text{stab}_m \cap \text{stab}_n$. It follows that $\tau(mn) = \tau m \cdot \tau n = mn$, and therefore $\tau \in \text{stab}_{mn}$. We have shown that any two elements outside $\text{mem}(m) \cup \text{mem}(n)$ are in the same data orbit for mn . Since there is only one infinite data orbit, it follows that they are not in $\text{mem}(mn)$. ◀

5 Algebraic properties of data monoids

In this section we develop the algebraic theory of data monoids.

Local finiteness. Recall that a monoid is called *locally finite* if all of its finitely generated submonoids are finite. Much of Green theory works for locally finite monoids, which makes the following theorem important.

► **Theorem 6.** *Every orbit finite data monoid is locally finite.*

Preliminary work indicates that local finiteness holds under much weaker conditions than being orbit finite. Namely, for every language of data words that is defined in monadic second-order logic with order and data equality, the syntactic monoid is locally finite. This covers, for instance, languages recognized by nondeterministic register automata.

Green's relations. Suppose that M is a data monoid. Using the underlying monoid, one can define Green's relations on elements $m, n \in M$.

- $m \leq_{\mathcal{L}} n$ if $Mm \subseteq Mn$
- $m \leq_{\mathcal{R}} n$ if $mM \subseteq nM$
- $m \leq_{\mathcal{J}} n$ if $MmM \subseteq MnM$

We sometimes say that n is a *suffix* of m instead of writing $m \leq_{\mathcal{L}} n$. Likewise for *prefix* and *infix*, and the relations $\leq_{\mathcal{R}}$ and $\leq_{\mathcal{J}}$. Finally, we use the name *extension* for the converse of infix. Like in any monoid, these relations are transitive and reflexive, so one can talk about their induced equivalence relations $\sim_{\mathcal{L}}$, $\sim_{\mathcal{R}}$ and $\sim_{\mathcal{J}}$. Equivalence classes of these relations are called \mathcal{L} -classes, \mathcal{R} -classes and \mathcal{J} -classes. An \mathcal{H} -class is defined to be the intersection of an \mathcal{L} -class and an \mathcal{R} -class.

In data monoids, we add a new relation, by lifting the order $\leq_{\mathcal{J}}$ to orbits. One could do the same for \mathcal{R} and \mathcal{L} , but in our application of data monoids to characterize of first-order logic, we only use the case for \mathcal{J} . We write $m \leq_{\mathcal{O}} n$ if there is some data renaming τ such that $\tau m \leq_{\mathcal{J}} n$. The letter \mathcal{O} stands for orbit. In terms of ideals, $\tau(MmM) \subseteq MnM$, or equivalently $M \cdot \tau m \cdot M \subseteq MnM$, or also equivalently $m \in M \cdot \tau n \cdot M$.

► **Lemma 7.** *In any data monoid, the relation $\leq_{\mathcal{O}}$ is transitive.*

Proof. Suppose that $x \geq_{\mathcal{O}} y \geq_{\mathcal{O}} z$ holds. This means that z can be written as $\tau(pyq)$ and y can be written as $\sigma(mxn)$, for some data renamings τ, σ and data monoid elements p, q, m, n . Combining the two, we get

$$z = \tau(p \cdot \sigma(mxn) \cdot q) = \tau(p \cdot \sigma m \cdot x \cdot \sigma n \cdot q) \in M \cdot \tau x \cdot M.$$

◀

Of course, in an orbit finite data monoid there are finitely many \mathcal{O} -classes, since each is a union of orbits.

\mathcal{J} -classes. We present a version of the Memory Theorem for \mathcal{J} -classes. For a \mathcal{J} -class J , we define its memory as

$$\text{mem}(J) = \bigcap_{m \in J} \text{mem}(m).$$

Clearly this set is finite, as an intersection of finite sets.

► **Theorem 8 (Memory Theorem for \mathcal{J} -classes).** *Let J be a \mathcal{J} -class. Any data renaming τ that is the identity on $\text{mem}(J)$ satisfies $\tau J = J$.*

Quotient under a \mathcal{J} -class. Suppose that J is a \mathcal{J} -class in a data monoid. Let \simeq_J be the equivalence relation on the data monoid defined by:

$$m \simeq_J n \quad \text{if } m = n \text{ or neither of } m, n \text{ is an infix of some element of } J.$$

► **Lemma 9.** *The relation \simeq_J is a congruence of data monoids, assuming the constants are $\text{mem}(J)$. Consequently, the quotient function denoted $h_{/J}$, is a data monoid homomorphism, with constants $\text{mem}(J)$.*

Orbit-equivalent \mathcal{J} -classes. We say that two \mathcal{J} -classes $J, K \subseteq M$ of a data monoid are *orbit-equivalent* if there is some data renaming τ such that $\tau J = K$.

► **Lemma 10.** *Orbit-equivalent \mathcal{J} -classes form antichains in the order $\leq_{\mathcal{J}}$.*

6 First-order logic

We test the algebra on first-order logic, yielding the main result of the paper, Theorem 11. We extend first-order logic, as described in the introduction, by a unary predicate $d(x)$ for every data value d , which selects positions with data value d . A predicate $d(x)$ is called a *data predicate*. Any formula uses a finite number of data predicates, and its language is invariant under data renamings that preserve the data values of the data predicates it uses. The formulas produced below will use data predicates only for the constants.

► **Theorem 11.** *Let L be a data language whose syntactic data monoid M_L is orbit finite. Then L is definable in first-order logic if and only if M_L is aperiodic.*

The implication from first-order definable to aperiodic is proved in the same way as usual, e.g. using an Ehrenfeucht-Fraïssé game. We focus on the more difficult implication. The proof is an induction, which needs a more detailed statement, as presented below.

► **Proposition 12.** *Let L be a data language recognized by a data monoid homomorphism $h : D^* \rightarrow M$ into a data monoid that is orbit finite and aperiodic. For every $m \in M$, the language $h^{-1}(m)$ can be defined by a sentence φ_m of FO that uses data predicates for data values in $\text{mem}(m)$.*

Note that in Theorem 11 and Proposition 12, there is an implicit set of constants $C \subseteq D$, which influences the definitions of data language, data monoid, and homomorphism.

Before we prove Proposition 12, we show how it implies Theorem 11. The language L is a union of orbits. Let m_1, \dots, m_n be chosen representatives of these orbits. We know that L is the same as the union

$$\Gamma_C \cdot m_1 \cup \dots \cup \Gamma_C \cdot m_n.$$

Each of the languages in the finite union above can be defined in first-order logic, thanks to Proposition 12 and the following lemma.

► **Lemma 13.** *Suppose that L is a language defined by a formula of first-order logic with data predicates A . For any $B \subseteq A$, the language*

$$\Gamma_B L = \{\tau w : \tau \in \Gamma_B, w \in L\}$$

is defined by a formula of first-order logic with data predicates B .

We now proceed with the proof of Proposition 12. The proof is by induction on two parameters, ordered lexicographically. The first, and more important, parameter is the number of orbits in the data monoid M . The second parameter is the position of the \mathcal{O} -class of m in the (inverse of the) order $\leq_{\mathcal{O}}$. We begin with the \mathcal{O} -class of the empty word, and proceed to bigger words, the idea being that it is easier to write formulas for infixes than extensions.

The induction base is done the same way as the induction step, so we omit it, and only do the induction step. Let O be the \mathcal{O} -class of m . By induction assumption, we can use formulas for elements with strictly bigger \mathcal{O} -classes, which correspond to infixes of m . There are two cases to consider, depending on whether O is a \mathcal{J} -class, or not. The case when O is a \mathcal{J} -class is more similar to the proof for finite alphabets.

O is not a \mathcal{J} -class. Let $J \subsetneq O$ be the \mathcal{J} -class that contains m . Let us expand the set of constants from C to $\text{mem}(J)$. Consider the data monoid homomorphism

$$h_{/J} : M \rightarrow M_{/J}$$

which squashes all elements that are not infixes of J into a single element, as in Lemma 9. Recall the data monoid homomorphism h that recognizes L . It is easy to see that the composition $g = h_J \circ h$ recognizes $h^{-1}(m)$, since g maps the same elements to m as h .

► **Lemma 14.** *$M_{/J}$ has fewer orbits than M .*

Thanks to the above lemma, we can use the induction assumption to produce the formula for m . This finishes the case when O is not a \mathcal{J} -class.

O is a \mathcal{J} -class. For $n \in M$, define K_n to be the set of data words that have image n under h , but all of their proper infixes have image in a strictly bigger \mathcal{O} -class than O . Using the induction assumption, we prove the following lemma.

► **Lemma 15.** *For each $n \in M$, the language K_n can be defined by a formula of first-order logic with data predicates $\text{mem}(n)$.*

► **Lemma 16.** *There are finitely many elements m_1, \dots, m_k and a subset $B \subseteq \text{mem}(m)$ such that for every $n \in M$,*

$$n \sim_{\mathcal{R}} m \quad \text{iff} \quad n \in \Gamma_B\{m_1, \dots, m_k\}.$$

By combining Lemmas 15, 13 and 16, we get the following lemma.

► **Lemma 17.** *The infinite union of languages K_n , ranging over $n \sim_{\mathcal{R}} m$, can be defined by a formula of first-order logic with data predicates $\text{mem}(m)$.*

Using Lemma 17 and its symmetric variant for \mathcal{L} -classes, we can write a formula of first-order logic, with data predicates $\text{mem}(m)$, which describes the set, call it L_m , of elements that have a prefix in the \mathcal{R} -class of m , and a suffix in the \mathcal{L} -class of m :

$$L_m = \bigcup_{n \sim_{\mathcal{R}} m} K_n D^* \quad \cap \quad \bigcup_{n \sim_{\mathcal{L}} m} D^* K_n.$$

If a word w belongs to L_m , then its image $n = h(w)$ satisfies $n \leq_{\mathcal{R}} m$, likewise for $\leq_{\mathcal{L}}$. By the following lemma, if we additionally assume that $n \sim_{\mathcal{J}} m$, then also $n \sim_{\mathcal{L}} m$ and $n \sim_{\mathcal{R}} m$, and hence $n \sim_{\mathcal{H}} m$.

► **Lemma 18.** *In a locally finite monoid, $m \sim_{\mathcal{J}} n$ and $m \leq_{\mathcal{R}} n$ imply $m \sim_{\mathcal{R}} n$. Likewise for \mathcal{L} .*

In other words, if a word belongs to L_m and has the same \mathcal{J} -class as m , then its image is in the \mathcal{H} -class of m . In a locally finite aperiodic monoid, this means that its image is simply m :

► **Lemma 19.** *In a locally finite aperiodic monoid, all \mathcal{H} -classes are singletons.*

Above we have shown that a word in L_m has image m , using the assumption that its image is in the same \mathcal{J} -class as m . Therefore, a word has image m if and only if it belongs to language L_m , which is definable in first-order logic, and its \mathcal{J} -class is O . Let K be the set of data words that have an infix in O , but are not in O . As argued before, a word has image m if and only if it belongs to the difference $L_m - K$. Therefore, to conclude we only need the following lemma.

► **Lemma 20.** *The language K can be defined in first-order logic.*

Proof. When does a word w belong to K ? Consider the smallest infix of w that belongs to K . Cutting off the first or last letter of that infix leads to a word in O . Therefore:

$$K = D^* \cdot \left(\bigcup_{\substack{d \in D, n \in O \\ h(d) \cdot n \notin O}} d \cdot h^{-1}(n) \cup \bigcup_{\substack{d \in D, n \in O \\ n \cdot h(d) \notin O}} h^{-1}(n) \cdot d \right) \cdot D^*$$

The important observation is that it does no harm to replace $h^{-1}(n)$ by the language L_n defined above. This is because L_n contains $h^{-1}(n)$ and is contained in $h^{-1}(n) \cup K$. Therefore,

$$K = D^* \cdot \left(\bigcup_{\substack{d \in D, n \in O \\ h(d) \cdot n \notin O}} d \cdot L_n \cup \bigcup_{\substack{d \in D, n \in O \\ n \cdot h(d) \notin O}} L_n \cdot d \right) \cdot D^*$$

We deal with the infinite union using Lemma 13. ◀

7 Further work

Characterize other logics. It is natural to extend the characterization of first-order logic to other logics. Candidates that come to mind include first-order logic with two variables, or various logics inspired by XPath, or piecewise testable languages. Also, it would be interesting to see the expressive power of languages recognized by orbit-finite data monoids. This class of languages is incomparable in expressive power to first-order logic, e.g. the first-order definable language “some data value appears twice” is not recognized by an orbit-finite data monoid. It would be nice to see a logic, maybe a variant of monadic second-order logic, with the same expressive power as orbit-finite data monoids.

More structure on the data values. In this paper, we study languages that are closed under arbitrary renamings of data values. One could be interested in weaker requirements, that give more general language classes. For instance, consider languages $L \subseteq \mathbb{R}^*$ that are closed under order-preserving renamings. Another, very interesting example, concerns languages of timed automata.

Use mechanisms more powerful than monoids. Even if we are interested in languages that are preserved under arbitrary data renamings, orbit finite data monoids have weak expressive power. Contrary to the case of finite alphabets, over infinite alphabets, (orbit

finite) data monoids are strictly less expressive than the natural automaton counterpart⁵. If we only require the syntactic left-to-right automaton to be orbit finite, we get a larger class of languages. This larger class includes the language “the first letter in the word appears also on some other position”, which has a syntactic monoid with infinitely many orbits. Therefore, one can ask: is it decidable if an orbit finite automaton recognizes a language that can be defined in first-order logic? We conjecture that this problem is decidable, and even that a necessary and sufficient condition is aperiodicity of the syntactic monoid (which need not be orbit finite). Aperiodicity is not, in general, the right condition for first-order logic, as witnessed by the language “words with an even number of distinct letters”, which has an aperiodic syntactic monoid (sic!), but is not definable in first-order logic.

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⁵ This counterpart has the same expressive power as deterministic memory automata of Francez and Kaminski [6]