



HAL
open science

A dichotomic Newton-Puiseux algorithm using dynamic evaluation

Adrien Poteaux, Martin Weimann

► **To cite this version:**

Adrien Poteaux, Martin Weimann. A dichotomic Newton-Puiseux algorithm using dynamic evaluation. 2017. hal-01578214v2

HAL Id: hal-01578214

<https://inria.hal.science/hal-01578214v2>

Preprint submitted on 3 Jan 2018 (v2), last revised 3 Dec 2018 (v3)

HAL is a multi-disciplinary open access archive for the deposit and dissemination of scientific research documents, whether they are published or not. The documents may come from teaching and research institutions in France or abroad, or from public or private research centers.

L'archive ouverte pluridisciplinaire **HAL**, est destinée au dépôt et à la diffusion de documents scientifiques de niveau recherche, publiés ou non, émanant des établissements d'enseignement et de recherche français ou étrangers, des laboratoires publics ou privés.



Distributed under a Creative Commons Attribution - NonCommercial - ShareAlike 4.0 International License

A dichotomic Newton-Puiseux algorithm using dynamic evaluation

Adrien Poteaux,
CRISAL
Université de Lille 1
UMR CNRS 9189
Bâtiment M3
59655 Villeneuve d'Ascq, France
adrien.poteaux@univ-lille1.fr

Martin Weimann
GAATI
Université de Polynésie Française
BP 6570
98702 Faa'a
martin.weimann@upf.pf

January 3, 2018

Let $F \in \mathbb{K}[X, Y]$ be a polynomial of total degree D defined over a field \mathbb{K} of characteristic zero or greater than D . Assuming F separable with respect to Y , we provide an algorithm that computes all rational Puiseux expansions of F above $X = 0$ in less than $\mathcal{O}(Dv)$ operations in \mathbb{K} , where v is the valuation of the resultant of F and its partial derivative with respect to Y . To this aim, we use a divide and conquer strategy and replace univariate factorization by dynamic evaluation. As a first main corollary, we compute the irreducible factors of F in $\mathbb{K}[[X]][Y]$ up to an arbitrary precision X^N with $\mathcal{O}(D(v + N))$ arithmetic operations. As a second main corollary, we compute the genus of the plane curve defined by F with $\mathcal{O}(D^3)$ arithmetic operations and, if $\mathbb{K} = \mathbb{Q}$, with $\mathcal{O}(hD^3)$ bit operations where h is the logarithmic height of F .

1 Introduction.

Let $F \in \mathbb{K}[X, Y]$ be a bivariate polynomial defined over a field \mathbb{K} . Assuming that the characteristic p of \mathbb{K} is zero or large enough, it is well known that for any $x_0 \in \overline{\mathbb{K}}$, the roots of F (considered as a univariate polynomial in Y) may be expressed as fractional Laurent power series in $(X - x_0)$ with coefficients in $\overline{\mathbb{K}}$. These are the (classical) *Puiseux series*¹ of F above x_0 . Puiseux series are fundamental objects of the theory of algebraic curves [32, 6] and provide important information. The reader will find many applications in previous papers [28, 27] by the first author.

In [12, 13], D. Duval introduced *rational Puiseux expansions*. They are more convenient to compute and provide more geometrical informations (see Section 2). Our first contribution concerns the computations of *singular parts* of rational Puiseux expansions.

In the sequel, we denote d_X and d_Y the partial degrees of F in respectively X and Y , and D its total degree. We also let $v = v_X(R_F)$ stand for the X -valuation of the resultant $R_F := \text{Res}_Y(F, F_Y)$ of F and its Y -derivative.

Theorem 1. *Let $F \in \mathbb{K}[X][Y]$ be a primitive and separable polynomial in Y . Assume that $p = 0$ or $p > d_Y$. There exists an algorithm² that computes singular parts of Puiseux series of F above $x_0 = 0$ in less than $\mathcal{O}(d_Y v)$ arithmetic operations over \mathbb{K} .*

This improves results of [28] (the complexity bound therein is in $\mathcal{O}(\rho v d_Y)$ where ρ is the number of rational Puiseux expansions - bounded by d_Y). From that we will deduce:

¹terms written *in italics* in this introduction will be defined in Section 2 or 5.1.

²note that our main algorithms are Las Vegas; this is only due to the computation of primitive elements; see Sections 3.1 and 5.2

Theorem 2. *Let F as in Theorem 1. There exists an algorithm that computes the singular part of Puiseux series of F above all critical points in less than $\mathcal{O}(d_Y^2 d_X) \subset \mathcal{O}(D^3)$ arithmetic operations.*

This result improves the bound in $\mathcal{O}(d_Y^2 d_X^3) \subset \mathcal{O}(D^5)$ of [25, 26]; note also that Proposition 12 of [28] suggested a way to get a bound in $\mathcal{O}(d_Y^3 d_X) \subset \mathcal{O}(D^4)$.

Using the Riemann-Hurwitz formula, we easily deduce:

Corollary 1. *Assuming $p = 0$ or $p > D$, there exists an algorithm that computes the genus of a given geometrically irreducible algebraic plane curve over \mathbb{K} of degree D in less than $\mathcal{O}(D^3)$ arithmetic operations.*

Moreover, using the reduction criterion of [25, 27], we can bound the bit complexity of the genus computation (here $\text{ht}(P)$ stands for the maximum between the logarithm of the denominator of P , and the logarithm of the infinite norm of its numerator):

Corollary 2. *Let $\mathbb{K} = \mathbb{Q}(\gamma)$ be a number field, $0 < \epsilon \leq 1$ a real number and $F \in \mathbb{K}[X, Y]$. Denote M_γ the minimal polynomial of γ and w its degree. Then there exists a Monte Carlo algorithm that computes the genus of the curve $F(X, Y) = 0$ with probability of error less than ϵ and an expected number of word operations in:*

$$\mathcal{O}(d_Y^2 d_X w^2 \log^2 \epsilon^{-1} [\text{ht}(M_\gamma) + \text{ht}(F)]).$$

With the same notations than in Corollary 2, we have:

Corollary 3. *Assuming that the degree of the square-free part of the resultant $\text{Res}_Y(F, F_Y)$ is known, there exists a Las Vegas algorithm that computes the genus of the curve $F(X, Y) = 0$ with an expected number of word operations in:*

$$\mathcal{O}(d_Y^2 d_X w^2 [\text{ht}(M_\gamma) + \text{ht}(F)]).$$

Finally, as described below, our algorithm will induce a fast analytic factorisation of the polynomial F . More precisely, one gets the following result:

Theorem 3. *Under hypothesis of Theorem 1, there exists an algorithm that computes the irreducible analytic factors of F in $\mathbb{K}[[X]][Y]$ with precision $N \in \mathbb{N}$ in less than $\mathcal{O}(d_Y(v + N))$ - thus $\mathcal{O}(d_Y N)$ if $N \geq v$ - arithmetic operations in \mathbb{K} , plus the cost of one univariate factorisation of degree at most d_Y .*

Theorem 3 has a particular interest with regards to factorization in $\mathbb{K}[X, Y]$ or $\overline{\mathbb{K}}[X, Y]$: when working along a critical fiber, one can take advantage of some combinatorial constraints imposed by ramification when recombining analytic factors into rational factors (see [33]). In particular, if F is irreducible in $\mathbb{K}[[X]][Y]$, then it is in $\mathbb{K}[X, Y]$.

Main ideas. We now describe the main ingredients of our algorithm to improve the results of [28].

Idea 1. We concentrate on the monic case. The roots above $(0, \infty)$ require special care, explaining that the results depend on $v = v_X(\text{Res}_Y(F, F_Y))$ and not on $v_X(\text{Disc}_Y(F)) \leq v$. However, equality holds when F is monic.

Idea 2. The key idea is to use tight truncation bounds for the powers of X in the course of the algorithm. Even if the bound $n = v$ can be reached for *some* Puiseux series, it is actually always possible to compute at least half of them using a bound $n \in \mathcal{O}(v/d_Y)$ (see Section 3). We will use such an idea to show that we can either prove that F is irreducible (and compute its Puiseux series), either compute a factorisation $F = GH \pmod{X^n}$ with $n \in \mathcal{O}(v/d_Y)$ and $d_Y(H) \leq d_Y/2$; and that this can be done in the aimed bound of Theorem 2.

Idea 3. The fiber $X = 0$ being critical, the polynomials $G(0, Y)$ and $H(0, Y)$ may not be coprime. Therefore, we cannot use the classical Hensel lemma to lift the factorisation of the previous point. We will show (see Lemma 10) that it is possible to adapt the Hensel lemma to our case.

Idea 4. Nevertheless, such a result might require to start with a factorisation $F = G \cdot H \pmod{X^k}$ with $k \in \Theta(v)$. We prove in Section 4 that in our context $k \in \mathcal{O}(v/d_Y)$.

Idea 5. We then apply this divide and conquer strategy until we get an analytic factorisation of $F \bmod X^v$, together with the singular parts of its Puiseux series.

Idea 6. Finally, as suggested in the conclusion of [28], computing univariate factorisations needed in the *Newton-Puiseux algorithm* can be costly, in particular when we consider the computation of singular parts of Puiseux series above *all* critical points. As suggested in [28], we will use dynamic evaluation [11, 10] to avoid this bottleneck, which leads us to work over product of fields. As in [13], this will also enable the results to stand over characteristic 0 fields.

Organisation of the paper. We remember the classical definitions related to Puiseux series and the description of the *rational Newton Puiseux algorithm* of [13] in Section 2. The remaining of the paper will be divided into two parts:

1. We first assume the input polynomial to be monic and defined above a field. As in [26, 28], we will use univariate factorisation. Sections 3 and 4 will present Ideas 2 to 5:
 - Section 3 presents a variant of the rational Newton-Puiseux algorithm. It enables us to use tight truncation bounds. In particular, we get half of the Puiseux series above $x_0 = 0$ in $\mathcal{O}(d_Y v)$ arithmetic operations, plus the cost of univariate factorisations.
 - Given these Puiseux series, we show in Section 4 how to compute two factors of $F: G$ corresponding to the Puiseux series computed so far, and H satisfying $d_Y(H) \leq d_Y/2$.
 - Then, we recursively compute the Puiseux series of H . This defines a divide and conquer algorithm to compute all singular parts of Puiseux series above $x_0 = 0$ in less than $\mathcal{O}(d_Y v)$ operations (plus univariate factorisations once again).

Finally, we will explain in Section 4.5 how to deal with the non monic case as sketched, detailing how to deal with roots above $(0, \infty)$.

2. The next step is to get rid of univariate factorisations. They are costly when the characteristic is 0. Even above finite fields, it has been mentioned in [28] that we cannot to get a bound depending on v . This is a problem when considering the computation of Puiseux series above *all* critical points. Avoiding them is Idea 6, i.e. dynamic evaluation. In this context, we are led to work over non integral rings, and we have to pay attention to several points. For instance, the computation of the Newton polygon is not straightforward: we first need to check if the coefficients of the input polynomials are *regular* (not zero divisors); if not, we have to perform some suitable splittings. We detail this in Section 5, where we prove Theorem 1 using results of [10].

Finally, Section 6 deals with the computation of the desingularisation of the curve above all its critical points. Combining together the previous results, we get a complexity bound, as good as - up to logarithmic factors - the best known algorithm to compute bivariate resultants. This is Theorem 2 and associated corollaries. We then prove Theorem 3 in Section 7.

A brief state of the art. In [13], D. Duval defines the rational Newton-Puiseux algorithm over a field \mathbb{K} with characteristic 0. The complexity analysis therein provides a number of arithmetic operations in less than $\mathcal{O}(d_Y^6 d_X^2)$ when F is monic (without using any fast algorithm). Note that this algorithm does not compute any univariate factorisation (it uses the D5-principle instead), and can trivially be generalised when $p > d_Y$.

In [25, 26], an algorithm with complexity $\mathcal{O}(d_Y v^2 + d_Y v \log(p^c))$ is provided over $\mathbb{K} = \mathbb{F}_{p^c}$, with $p > d_Y$. Moreover, from this bound, one gets an algorithm that computes the singular parts of Puiseux series of a polynomial $F \in \mathbb{K}[X, Y]$ above *all* critical points in $\mathcal{O}(d_Y^3 d_X^2 \log(p^c))$. In [28], with the same assumption on the base field, an algorithm is given to compute the singular part of Puiseux series over $x_0 = 0$ in less than $\mathcal{O}(\rho d_Y v + \rho d_Y \log(p^c))$ arithmetic operations. These two algorithms rely on univariate factorisation over finite fields. Therefore, they cannot be directly extended to the 0 characteristic case. This also explains why the second result does not provide an improved bound for the computation of Puiseux series above *all* critical points.

There are other methods to compute Puiseux series or analytic factorisation, as generalised Hensel constructions [17], or the Montes algorithm [23, 2] (which works over general local fields). Several of these methods and a few others have been commented on in previous papers by the first author [27, 28]. Also, there exist algorithms for the genus based on linear differential operators and avoiding the computation of Puiseux series [21, 4]. To our knowledge, none of these methods have been proved to provide a complexity which fits in the bounds obtained in this paper.

Acknowledgment. This paper is dedicated to Marc Rybowicz, who passed away in November 2016 [14]. The first ideas of this paper actually came from a collaboration between Marc and the first author in the beginning of 2012, that led to [28] as a first step towards the divide and conquer algorithm presented here. We also thank Francois Lemaire for many useful discussions we had on dynamic evaluation.

2 Main definitions and classical algorithms.

2.1 Puiseux series.

Denote \mathbb{K} a field and consider $F \in \mathbb{K}[X, Y]$ as in Section 1. Up to a change of variable $X \leftarrow X + x_0$, it is sufficient to give definitions and properties for the case $x_0 = 0$. Under the assumption that $p = 0$ or $p > d_Y$, the well known Puiseux theorem asserts that the d_Y roots of F (viewed as a univariate polynomial in Y) lie in the field of Puiseux series $\cup_{e \in \mathbb{N}} \overline{\mathbb{K}}((X^{1/e}))$. See [6, 15, 32] or most textbooks about algebraic functions for the 0 characteristic case. When $p > d_Y$, see [9, Chap. IV, Sec. 6]. It happens that these Puiseux series can be grouped according to the field extension they define. Following Duval [13, Theorem 2], we consider decompositions into irreducible elements:

$$\begin{aligned} F &= \prod_{i=1}^{\rho} F_i \text{ with } F_i \text{ irreducible in } \mathbb{K}[[X]][Y] \\ F_i &= \prod_{j=1}^{f_i} F_{ij} \text{ with } F_{ij} \text{ irreducible in } \overline{\mathbb{K}}[[X]][Y] \\ F_{ij} &= \prod_{k=0}^{e_i-1} \left(Y - S_{ij}(X^{1/e_i} \zeta_{e_i}^k) \right) \text{ with } S_{ij} \in \overline{\mathbb{K}}((X)) \end{aligned}$$

where ζ_{e_i} is a primitive e_i -th root of unity in $\overline{\mathbb{K}}$. Primitive roots are chosen so that $\zeta_{ab}^b = \zeta_a$.

Definition 1. The d_Y fractional Laurent series $S_{ijk}(X) = S_{ij}(X^{1/e_i} \zeta_{e_i}^k) \in \overline{\mathbb{K}}((X^{1/e_i}))$ are called the *classical Puiseux series* of F above 0. $e_i \in \mathbb{N}$ is the *ramification index* of S_{ij} . If $S_{ij} \in \overline{\mathbb{L}}[[X^{1/e_i}]]$, we say that S_{ij} is defined at $X = 0$.

Proposition 1. The $\{F_{ij}\}_{1 \leq j \leq f_i}$ have coefficients in a degree f_i extension \mathbb{K}_i of \mathbb{K} . They are conjugated by the action of the Galois group of \mathbb{K}_i/\mathbb{K} . We call \mathbb{K}_i the residue field of R_i and f_i its residual degree. We have $\sum_{i=1}^{\rho} e_i f_i = d_Y$.

Proof. First claim is [13, Section 1]. For the second one, see e.g. [9, Chapter 4, Section 1]. \square

This leads to the definition of rational Puiseux expansions (note that it is possible to construct classical Puiseux series from a system of rational Puiseux expansions - see e.g. [28, Section 2]):

Definition 2. A system of *rational Puiseux expansions* over \mathbb{K} (\mathbb{K} -RPE) of F above 0 is a set $\{R_i\}_{1 \leq i \leq \rho}$ such that:

- $R_i(T) \in \mathbb{K}_i((T))^2$;
- $R_i(T) = (X_i(T), Y_i(T)) = (\gamma_i T^{e_i}, \sum_{l=-n_i}^{\infty} \beta_{il} T^l)$, with $n_i \in \mathbb{Z}$, $\gamma_i \neq 0$ and $\beta_{i, n_i} \neq 0$;
- R_i is a parametrisation of F_i , i.e. $F_i(X_i(T), Y_i(T)) = 0$;
- the parametrisation is irreducible, i.e. e_i is minimal.

We call $(X_i(0), Y_i(0))$ the *center* of R_i . We have $Y_i(0) = \infty$ if $n_i < 0$, which happens only for non monic polynomials.

Throughout this paper, we will truncate the powers of X of polynomials or series. To that purpose, we introduce the following notation: given $\tau \in \mathbb{Q}$ and a Puiseux series S with ramification index e , we denote $[S]^\tau = \sum_{k \leq N} \alpha_k X^{k/e}$ where $N = \max\{k \in \mathbb{N} \mid \frac{k}{e} \leq \tau\}$. We generalize this notation to elements of $\overline{\mathbb{K}}((X^{1/e}))[Y]$ by applying it coefficient-wise. In particular, if $H \in \mathbb{K}[[X]][Y]$ is defined as $H = \sum_i (\sum_{k \geq 0} h_{ik} X^k) Y^i$, then $[H]^\tau = \sum_i (\sum_{k=0}^{\lfloor \tau \rfloor} h_{ik} X^k) Y^i$.

Definition 3. The *regularity index* r of a Puiseux series S of F is the least integer $N \geq \min(0, e v_X(S))$ such that, if $\lceil S \rceil^{\frac{N}{e}} = \lceil S' \rceil^{\frac{N}{e}}$ for some Puiseux series S' of F , then $S = S'$. We call $\lceil S \rceil^{\frac{r}{e}}$ the *singular part* of S in F .

Roughly speaking, the regularity index is the number of terms necessary to “separate” a Puiseux series from all the others (with a special care when $v_X(S) < 0$). Since regularity indices of all Puiseux series corresponding to the same rational Puiseux expansion are equals, we define:

Definition 4. The *singular part* of a rational Puiseux expansion R_i of F is the pair

$$\left(\gamma_i T^{e_i}, \Gamma(T) = \sum_{k=n_i}^{r_i} \beta_{ik} T^k \right),$$

where r_i is the regularity index of R_i , i.e. the one of any Puiseux series associated to R_i .

Once such singular part has been computed, the Implicit Function Theorem ensures us that one can compute the series up to an arbitrary precision. This can be done in quasi linear time by using a Newton operator [18, Corollaries 5.1 and 5.2, page 251].

Finally, we formalise the *precision* of Puiseux series and rational Puiseux expansions that we will use in our algorithms. We also define the notion of pseudo-parametrisation:

Definition 5. We say that

- $S' \in \overline{\mathbb{K}}((X^{1/e}))$ is a Puiseux series of F known with precision n if there exists a Puiseux series S of F s.t. $\lceil S' \rceil^n = \lceil S \rceil^n$,
- $R' = (\gamma' T^{e'}, \Gamma'(T)) \in \mathbb{K}((T))^2$ is a RPE of F known with precision n if there exists a RPE $R = (\gamma T^e, \Gamma(T))$ of F such that e' divides e , γ' divides γ and $\lceil \Gamma'((X/\gamma')^{1/e'}) \rceil^n = \lceil \Gamma((X/\gamma)^{1/e}) \rceil^n$ for some choices of roots of unity,
- $\pi = (\gamma X^e, \Gamma(X) + \alpha X^\tau Y)$ is a pseudo-parametrisation of F if $\pi(T, 0)$ is a RPE of F known with precision $\frac{\tau}{e}$.

2.2 The rational Newton Puiseux algorithm.

Our algorithm in Section 3 is a variant of the well known Newton-Puiseux algorithm [32, 6]. We now explain (roughly speaking) the idea of this algorithm following an example, and then describe the variant of D. Duval [13, section 4], since we will use the improvements therein.

Let $F_0(X, Y) = Y^6 + Y^5 X + 5 Y^4 X^3 - 2 Y^4 X + 4 Y^2 X^2 + X^5 - 3 X^4$ and consider its Puiseux series computation. According to the well known Puiseux theorem (see e.g. [6, Section 8.3]), the first term of any such series $S(X)$ is $\alpha X^{\frac{m}{q}}$ for some $\alpha \in \overline{\mathbb{K}}$ and $(m, q) \in \mathbb{N}^2$.

We have $F_0(X, \alpha X^{\frac{m}{q}} + \dots) = \alpha^6 X^{\frac{6m}{q}} + \alpha^5 X^{\frac{5m}{q}+1} + 5 \alpha^4 X^{\frac{4m}{q}+3} - 2 \alpha^4 X^{\frac{4m}{q}+1} + 4 \alpha^2 X^{\frac{2m}{q}+2} + X^5 - 3 X^4 + \dots$. To get $F_0(X, S(X)) = 0$, at least two terms of the previous sum must cancel one another, i.e. (m, q) must be chosen so that two or more of the exponents coincide. To that purpose, we use the following definition:

Definition 6. Given $F(X, Y) = \sum_{i,j} f_{ij} X^j Y^i$, $\text{Supp}(F) = \{(i, j) \in \mathbb{N}^2 \mid f_{ij} \neq 0\}$ is the *support* of F .

The condition on (m, q) can be translated as: two points of $\text{Supp}(F_0)$ belong to the same line $ma + qb = l$. In order to increase the X -order of the evaluation, there must be no point under this line. In this example, we have two such lines: $a + 2b = 6$ and $a + b = 4$. They define the Newton polygon of F_0 :

Definition 7. The *Newton polygon* $\mathcal{N}(F)$ of $F \in \mathbb{K}[X, Y]$ is the lower part of the convex hull of its support.

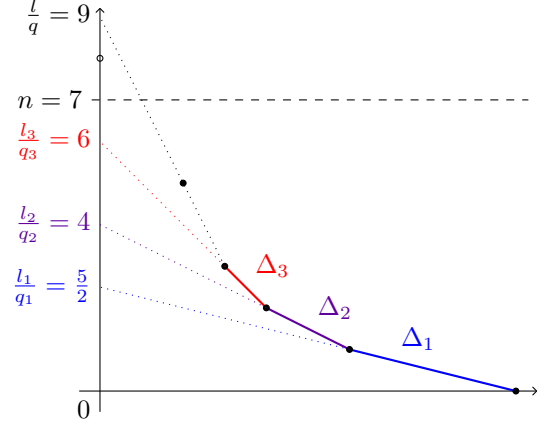
Remark 1. In [28, Definition 6], a *modified* Newton polygon $\mathcal{N}^*(H)$ is defined, and used in the main algorithm. This is not necessary for our strategy, since we do not focus on getting *precisely* the singular part. But we will need it for the proof of Lemma 5 (see Remark 5), so we use it in the description of **RNPuiseux**.

In this paper, we will use low truncation bounds; in particular, we may truncate some points of the Newton polygon. In order to certify the correctness of the computed slopes, we will use the following definition.

Definition 8. Given $F \in \mathbb{K}[X, Y]$ and $n \in \mathbb{N}$, the n -truncated Newton polygon of F is the set $\mathcal{N}_n(F)$ composed of edges Δ of $\mathcal{N}([F]^n)$ that satisfy $\frac{l}{q} \leq n$ if Δ belongs to the line $ma + qb = l$. In particular, any edge of $\mathcal{N}_n(F)$ is an edge of $\mathcal{N}(F)$.

Example 1.

Let us consider $F_1 = Y^{10} + XY^6 + X^2Y^4 + X^3Y^3 + X^5Y^2 + X^8$ and $n = 7$; the picture here provides the truncated Newton polygon of F_1 with precision 7. Here we have $[F_1]^7 = Y^{10} + XY^6 + X^2Y^4 + X^3Y^3 + X^5Y^2$ and $\mathcal{N}([F_1]^7) = [(10, 0), (6, 1), (4, 2), (3, 3), (2, 5)]$. But the edge $[(3, 3), (2, 5)]$ is not part of $\mathcal{N}_n(F_1)$, as it belongs to $2a + b = 9$, and that there are points (i, j) so that $2i + j \leq 9$ and $j > 7$: from the knowledge of $[F_1]^7$, we cannot guaranty that $\mathcal{N}(F_1)$ contains an edge belonging to $2a + b = 9$. This is indeed wrong here, since $\mathcal{N}(F_1) = [(10, 0), (6, 1), (4, 2), (3, 3), (0, 8)]$.



We are now considering the choice of α corresponding to $a + 2b = 6$. We have $F_0(T^2, \alpha T) = (\alpha^6 - 2\alpha^4 + 4\alpha^2)T^6 - 3T^8 + \alpha^5T^7 + (5\alpha^4 + 1)T^{10} + \dots$, meaning that α must be a non zero root of $P(Z) = Z^6 - 2Z^4 + 4Z^2$.

Finally, to get more terms, we can apply recursively this strategy to the polynomial $F_0(X^q, X^m(Y + \alpha))$. Actually, it is more interesting to consider a root $\xi = \alpha^q$ of the polynomial $\phi(Z) = Z^2 - 2Z + 4$ (we have $P(Z) = Z^2\phi(Z^2)$ and are obviously not interested in the root $\alpha = 0$), which is the characteristic polynomial [13]:

Definition 9. If $F = \sum f_{ij}X^jY^i$, then the *characteristic polynomial* ϕ_Δ of $\Delta \in \mathcal{N}(F)$ is $\phi_\Delta(T) = \sum_{(a,b) \in \Delta} f_{ab}T^{\frac{a-a_0}{q}}$ where a_0 is the smallest value such that (a_0, b_0) belongs to Δ for some b_0 .

Up to now, we only described the ideas of the algorithm and provided all the necessary tools. We conclude this section by providing a more formal definition of the **RNPuisseux** algorithm for monic polynomials (see Section 4.4 for the non monic case); it uses two sub algorithms, for each we only provide specifications:

- **Bézout**, given $(q, m) \in \mathbb{Z}^2$ with $q > 0$, computes $(u, v) \in \mathbb{Z}^2$ s.t. $uq - mv = 1$ and $0 \leq v < q$
- **Factor** (\mathbb{K}, ϕ) computes the factorisation of ϕ over \mathbb{K} , given as a list of factors and multiplicities

Algorithm: **RNPuisseux** (F, \mathbb{K}, π)

Input: $F \in \mathbb{K}[X, Y]$ monic, \mathbb{K} a field and π the result of previous computations ($\pi = (X, Y)$ for the initial call)

Output: A set of singular parts of rational Puiseux expansions above $(0, 0)$ of F with its base field.

```

1  $\mathcal{R} \leftarrow \{\}$ ; // results of the algorithm will be grouped in  $\mathcal{R}$ 
2 foreach  $\Delta \in \mathcal{N}^*(F)$  do // we consider only negative slopes
3   Compute  $m, q, l, \phi_\Delta$  associated to  $\Delta$ ;
4    $(u, v) \leftarrow \text{Bézout}(m, q)$ ;
5   foreach  $(\phi, M)$  in Factor $(\phi_\Delta)$  do
6     Take  $\xi$  a new symbol satisfying  $\phi(\xi) = 0$ ;
7      $\pi_1 = \pi(\xi^v X^q, X^m(Y + \xi^u))$ ;
8     if  $M = 1$  then  $\mathcal{R} \leftarrow \mathcal{R} \cup \{(\pi_1(T, 0), \mathbb{K}(\xi))\}$ ;
9     else
10       $H(X, Y) \leftarrow F(\xi^v X^q, X^m(Y + \xi^u))/X^l$ ; // Puiseux transformation
11       $\mathcal{R} \leftarrow \mathcal{R} \cup \text{RNPuisseux}(H, \mathbb{K}(\xi), \pi_1)$ ;
12 return  $\mathcal{R}$ ;

```

The key improvement of this rational version is the distribution of ξ to both X and Y variables (line 10). This avoids to work with $\alpha = \xi^{1/q}$ and to introduce any useless field extension due to ramification (see [13, Section 4]).

2.3 Complexity notions.

We finally recall some classical complexity results, starting with the multiplication of univariate polynomials:

Definition 10. A (univariate) *multiplication time* is a map $M : \mathbb{N} \rightarrow \mathbb{R}$ such that:

- for any ring \mathbb{A} , polynomials of degree less than d in $\mathbb{A}[X]$ can be multiplied in at most $M(d)$ operations (multiplication or addition) in \mathbb{A} ;
- for any $0 < d \leq d'$, the inequalities $M(d) d' \leq M(d') d$ and $M(d d') \leq M(d) M(d')$ hold.

Lemma 1. *Let M be a multiplication time. Then we have:*

1. $M(d + d') \geq M(d) + M(d')$ for any $d, d' \in \mathbb{N}$,
2. $M(1) + M(2) + \dots + M(2^{k-1}) + M(2^k) \leq M(2^{k+1})$ for any $k \in \mathbb{N}$.

Proof. The first point is [16, Exercise 8.33]. The second one is an obvious consequence. \square

The best known multiplication time gives $M(d) \in \mathcal{O}(d \log(d) \log(\log(d))) \subset \mathcal{O}(d)$ [30, 7]. Note in particular that we do not have $M(d) M(d') \leq M(d d')$ but only $M(d) M(d') \leq M(d d') \log(d d')$. This is why we use Kronecker substitution.

Multiplication of multivariate polynomials. Consider two polynomials belonging to $\mathbb{A}[Z_1, \dots, Z_s]$. Denote d_i a bound for their degrees in Z_i . Then, by Kronecker substitution, they can be multiplied in less than $\mathcal{O}(M(2^{s-1} d_1 \dots d_s))$ operations in \mathbb{A} (it is straightforward to adapt [16, Corollary 8.28, page 247] to any number of variables). In particular, if s is constant, the complexity bound is $\mathcal{O}(M(d_1 \dots d_s))$.

Bivariate polynomials defined over an extension of \mathbb{K} . Let P be an irreducible polynomial over \mathbb{K} and denote $\mathbb{K}_P = \mathbb{K}[Z]/(P(Z))$. In Sections 3 and 4, we multiply two polynomials in $\mathbb{K}_P[X, Y]$ as follows: first perform the polynomial multiplication over $\mathbb{K}[X, Y, Z]$ as stated in the previous paragraph; then apply the reduction modulo P on each coefficient. Denote d_P the degree of P and d_X (resp. d_Y) a bound for the degree in X (resp. Y) of the considered polynomials. Then the total cost is $\mathcal{O}(M(d_X d_Y d_P))$ (see [16, Theorem 9.6, page 261] for the second point).

Finally, note that we postpone the discussion concerning the complexity of operations modulo triangular sets (needed for dynamic evaluation) in Section 5.2.

3 Refined truncation bounds to compute half of the Puiseux series.

In all this section, $F \in \mathbb{K}[X, Y]$ is monic and separable with respect to Y and the characteristic p of \mathbb{K} satisfies $p = 0$ or $p > d_Y$. Under such hypothesis, we will prove that we can compute at least half of the Puiseux series of F in less than $\mathcal{O}(d_Y v)$ arithmetic operations, not counting the factorisation of univariate polynomials.

Theorem 4. *There exists an algorithm that computes some RPEs R_1, \dots, R_λ of F known with precision at least $2v/d_Y$ and such that $\sum_{i=1}^\lambda e_i f_i \geq \frac{d_Y}{2}$. Not taking into account univariate factorisations, this can be done in less than $\mathcal{O}(M(v d_Y) \log(d_Y)) \subset \mathcal{O}(v d_Y)$ arithmetic operations over \mathbb{K} .*

Algorithm Half-RNP in Section 3.2 will be such an algorithm. It uses previous improvements by the first author and M. Rybowicz [25, 26, 28], and one additional idea, namely Idea 2 of Section 1, that we detail below.

3.1 Previous complexity improvements and Idea 2.

Given $P \in \mathbb{K}[Z]$, we denote \mathbb{K}_P the quotient $\mathbb{K}[Z]/(P(Z))$ and $d_P = d_Z(P)$. We need the following results of [26, 28].

Lemma 2. *Let $n \in \mathbb{N}$, $F \in \mathbb{K}_P[X, Y]$ and $\xi \in \mathbb{K}_P$ for some irreducible $P \in \mathbb{K}[Z]$. Denote Δ an edge of $\mathcal{N}(F)$ belonging to $ma + qb = l$, and $(u, v) = \text{Bézout}(m, q)$. One can reduce the computation of $F(\xi^v X^a, X^m(\xi^u + Y))/X^l$ modulo X^n to n univariate polynomial shifts over \mathbb{K}_P . This takes less than $\mathcal{O}(n M(d_Y d_P))$ arithmetic operations over \mathbb{K} .*

Proof. This is [26, Lemma 2, page 210]; Figure 1 illustrates the idea. Complexity also uses Kronecker substitution. \square

In [28], the number of recursive calls of the rational Newton-Puiseux algorithm is reduced from v to $\mathcal{O}(\rho \log(d_Y))$. This is due to what we call the *Abhyankar's trick* [1, Chapter 12].

Lemma 3. *Let $F = Y^{d_Y} + \sum_{i=0}^{d_Y-1} A_i(X) Y^i \in \mathbb{K}[X, Y]$. If the Newton polygon of $F(X, Y - A_{d_Y-1}/d_Y)$ has a unique edge $(\Delta) m a + q b = l$ with $q = 1$, then ϕ_Δ has several roots in $\overline{\mathbb{K}}$.*

In other words, after performing the shift $Y \leftarrow Y - A_{d_Y-1}/d_Y$, we can ensure that we will detect at least either a branch separation, a non integer slope, or a non trivial factor of the characteristic polynomial. This can happen no more than $\mathcal{O}(\rho \log(d_Y))$ times.

Lemma 4. *Let $F = Y^{d_Y} + \sum_{i=0}^{d_Y-1} A_i(X) Y^i \in \mathbb{K}_P[X, Y]$. One can compute $[F(X, Y - A_{d_Y-1}/d_Y)]^n$ in less than $\mathcal{O}(M(n d_Y(H) d_P))$ operations over \mathbb{K} .*

Proof. From our assumption on the characteristic of \mathbb{K} , this computation can be reduced to bivariate polynomial multiplication via [3, Problem 2.6, page 15]. The result follows (see Section 2.3). \square

We only need to apply this trick when the first term of *all* Puiseux series is the same. As in [28], we use in this section the following result, so that we perform such a bivariate shift only if necessary.

Proposition 2. *Let ϕ be in $\mathbb{K}_P[Y]$ and $d = \deg(\phi)$. There exists an algorithm **Oneroot** that, given \mathbb{K}_P and ϕ , decides if $\phi(Y) = c(Y - \xi)^d$ for some $c, \xi \in \mathbb{K}_P$. It takes less than $\mathcal{O}(M(d d_P))$ operations in \mathbb{K} .*

Proof. Writing $\phi(Y) = \sum_{i=0}^d \alpha_i Y^i$, it is sufficient to test whether the shifted polynomial $\phi(Y - \alpha_{d-1}/(d \alpha_d))$ is equal to $\alpha_d Y^d$; this can be done in $\mathcal{O}(M(d d_P))$ operations over \mathbb{K} via [3, Problem 2.6, page 15] and Kronecker substitution. \square

In order to provide the monicity assumption of Lemma 3, the well-known Weierstrass preparation theorem [1, Chapter 16] is used.

Proposition 3. *Let $G \in \mathbb{K}_P[X, Y]$ and assume it is not divisible by X . Then there exists unique \widehat{G} and U in $\mathbb{K}_P[[X]][Y]$ such that $G = \widehat{G}U$, with $U(0, 0) \neq 0$ and \widehat{G} a Weierstrass polynomial of degree $d_Y(\widehat{G}) = v_Y(G(0, Y))$. Moreover, RPEs of G and \widehat{G} centered at $(0, 0)$ are the same.*

The following result provides a complexity bound.

Proposition 4. *Let $G \in \mathbb{K}_P[X, Y]$ as in Proposition 3 and $n \in \mathbb{N}$. Denote \widehat{G} the Weierstrass polynomial of G . There exists an algorithm **WPT** that computes $[\widehat{G}]^n$ with $\mathcal{O}(M(n d_Y(G) d_P))$ operations in \mathbb{K} .*

Proof. This is [16, Theorem 15.18, page 451], using Kronecker substitution for multivariate polynomial multiplication. This theorem assumes that $\text{lc}_Y(F)$ is a unit, which is not necessarily the case here. However, formulae in [16, Algorithm 15.10, pages 445 and 446] can still be applied in our context: this is exactly [20, Algorithm Q, page 33]. \square

Finally, in both [26] and [28], we show that powers of X can be truncated modulo X^{v+1} during the algorithm while preserving the singular part of the Puiseux series. This explains the complexity bound of $\mathcal{O}(\rho v d_Y)$ of [28] by taking $n = v$ in the previous results.

Representation of residue fields. As explained in [26, Section 5.1], representing residue fields as multiple extensions can be costly. Therefore, we need to compute primitive representations each time we get a characteristic polynomial ϕ with degree 2 or more. Note that algorithms we use here are Las-Vegas (this is the only probabilistic part concerning our results on Puiseux series computation).

Proposition 5. *Let $P \in \mathbb{K}[Z]$ and $\phi \in \mathbb{K}_P[W]$ be two irreducible polynomials of respective degrees d_P and d_ϕ . Denote $d = d_P d_\phi$, and assume that there are at least d^2 elements in \mathbb{K} . There exists a Las-Vegas algorithm **Primitive** that computes an irreducible polynomial $P_1 \in \mathbb{K}[Z]$ with degree d together with an isomorphism $\Psi : \mathbb{K}_{P,\phi} \simeq \mathbb{K}_{P_1}$. It takes less than $\mathcal{O}(d^{\frac{\omega+1}{2}})$ arithmetic operations plus a constant number of irreducibility tests in $\mathbb{K}[Z]$ of degree at most d . Moreover, given $\alpha \in \mathbb{K}_{P,\phi}$, one can compute $\Psi(\alpha)$ with $\mathcal{O}(d_P M(d))$ operations over \mathbb{K} .*

Proof. Consider for instance [29, Section 2.2]; some details are provided in the proof of Proposition 15. \square

Remark 2. We do not precisely pay attention to the assumption about the number of elements in \mathbb{K} in this paper. Note that we will always have $d \leq d_Y$ in our context. Therefore, if \mathbb{K} is a finite field without enough elements, it is sufficient to build a degree 2 field extension since $p > d_Y$.

Idea 2: using lower truncation bounds. In this paper, we use tighter truncation bounds. We show that one can use a truncation bound $n \in \mathcal{O}(v/d_Y)$ and still get some informations, namely at least half of the singular parts of Puiseux series. Nevertheless, such a truncation bound will need to be updated in a different way than in [26, 28]. In particular, when there is a transform $X \leftarrow X^q$, one has to multiply the truncation bound by q . In the same way, one cannot divide the truncation bound by the degree t of the found extension anymore. Both these points are actually compensated by the use of the algorithm WPT. It will divide the degree in Y by the same amount, since it eliminates all the conjugates. Therefore, the size of the input polynomial will always be bounded by $\mathcal{O}(v)$ elements of \mathbb{K} . Tight truncations bounds (depending on the Puiseux series) are provided in Section 3.3; we also provide a bound depending on v .

3.2 The Half-RNP algorithm.

The following algorithm Half-RNP is a variant of algorithm ARNP of [28]. The main change is that we might compute only some of the RPEs (at least half of them for our applications). To this aim, we update the truncation bound in a slightly different way. Also, the computation of the output is presented in a different way (we eliminate the relations of [13, Section 4.1] in another way; this avoids to deal with field extensions twice), and we include how to deal with primitive elements for sake of completeness (see Proposition 5). Correctness of the output is proved in Section 3.3. We denote $v_i := v_X(H_Y(S_{i,j,k}(X)))$ for any Puiseux series $S_{i,j,k}$ associated to R_i .

Algorithm: Half-RNP(H, P, n, π)

Input: $P \in \mathbb{K}[Z]$ irreducible, $H \in \mathbb{K}_P[X, Y]$ separable and monic in Y with $d = d_Y(H) > 0$, $n \in \mathbb{N}$ (truncation order) and π the current pseudo-parametrisation ($P = Z$ and $\pi = (X, Y)$ for the initial call).

Output: all RPEs of R_i of H such that $n - v_i \geq r_i$, given with precision $(n - v_i)/e_i \geq r_i/e_i$.

```

1 if  $H = Y$  then return  $\pi(T, 0)$ ;
2 if  $\mathcal{N}_n(H) = [(0, \alpha), (d, 0)]$  with  $\alpha/d \in \mathbb{N}$  and  $\text{Oneroot}(\mathbb{K}_P, \phi_\Delta)$  then // Abhyankar's trick
3    $B \leftarrow A_{d-1}/d$ ;  $\pi_1 \leftarrow \lceil \pi(X, Y - B) \rceil^n$ ; //  $H = \sum_{i=0}^d A_i(X)Y^i$ 
4   if  $d = 1$  then return  $\pi_1(T, 0)$  else return Half-RNP( $\lceil H(X, Y - B) \rceil^n, P, n, \pi_1$ );
5  $\mathcal{R} \leftarrow \{\}$ ; // results will be grouped in  $\mathcal{R}$ 
6 foreach  $\Delta$  in  $\mathcal{N}_n(H)$  do //  $\Delta$  belongs to  $ma + qb = l$ ;  $u, v = \text{Bézout}(m, q)$ 
7   foreach  $(\phi, M)$  in Factor( $\mathbb{K}_P, \phi_\Delta$ ) do //  $\phi_\Delta$  the characteristic polynomial associated to  $\Delta$ 
8     if  $\deg_W(\phi) = 1$  then  $\xi, P_1, H_1, \pi_1 = -\phi(Z, 0), P, H, \pi$ ;
9     else // computing a primitive representation
10       $(P_1, \Psi) \leftarrow \text{Primitive}(P, \phi)$ ;
11       $\xi, H_1, \pi_1 \leftarrow \Psi(W), \Psi(H), \Psi(\pi)$ ; //  $\Psi : \mathbb{K}_{P, \phi} \rightarrow \mathbb{K}_{P_1}$  isomorphism
12       $\pi_2 \leftarrow \pi_1(\xi^v X^q, X^m(Y + \xi^u)) \bmod P_1$ ;
13       $H_2 \leftarrow \lceil H_1(\xi^v X^q, X^m(Y + \xi^u)) \rceil^{n_1} \bmod P_1$ ; //  $n_1 = qn - l$ 
14       $H_3 \leftarrow \text{WPT}(H_2, n_1)$ ;
15       $\mathcal{R} \leftarrow \mathcal{R} \cup \text{Half-RNP}(H_3, P_1, n_1, \pi_2)$ ; // no recursive call if  $n_1 < 0$ 
16 return  $\mathcal{R}$ ;

```

Remark 3. We have $d_X(\pi) \leq n e_i$ for any RPE deduced from π . This is obvious when π is defined from Line 3; changing X by X^q on Line 12 is also straightforward. Also, we have $m \leq n e_i$, since $\frac{m}{q} \leq \frac{l}{q} \leq n$ from Definition 8.

Remark 4. Except possibly at the first call, the polynomial H is then always a Weierstrass polynomial.

Theorem 4 is an immediate consequence of the following result, which will be proved in Section 3.4.

Theorem 5. *Let $F \in \mathbb{K}[X, Y]$ be as in Theorem 4. Then $\text{Half-RNP}(F, Z, 6v/d_Y, (X, Y))$ outputs a set of RPEs. Among them is a set R_1, \dots, R_λ known with precision at least $4v/d_Y$, satisfying moreover $v_i < 2v/d_Y$ and $\sum_{i=1}^\lambda e_i f_i \geq \frac{d_Y}{2}$. Not taking into account the cost of univariate factorisations, it takes less than $\mathcal{O}(M(v d_Y) \log(d_Y)) \subset \mathcal{O}(v d_Y)$ arithmetic operations over \mathbb{K} .*

3.3 Using tight truncations bounds.

We study here more carefully the RNPuiseux algorithm. This gives us an optimal truncation bound to compute a RPE of a monic polynomial F with this algorithm or Half-RNP . From this study, we also deduce an exact relation between v and this optimal bound. This section needs additional notations:

- F_Y is the derivative in Y of F and R_1, \dots, R_ρ are its rational Puiseux expansions.
- Let $m_{k,h} a + q_{k,h} b = l_{k,h}$, $1 \leq h \leq g_k$ be the successive edges encountered during the computation of the expansion R_k with RNPuiseux ; we denote $N_k = \sum_{h=1}^{g_k} \frac{l_{k,h}}{q_{k,1} \cdots q_{k,h}}$.
- For $1 \leq i \leq \rho$, r_i, e_i, f_i are the data associated to R_i . We let $v_i := v_X(F_Y(S))$, where S is any Puiseux series associated to R_i (this rational number does not depends on the choice of S).

Lemma 5. *With the above notations: for any $1 \leq i \leq \rho$, we have $N_i = \frac{r_i}{e_i} + v_i$.*

Proof. Denote $R_i = (\gamma_i X^{e_i}, \Gamma_i(X, Y))$ with $\Gamma_i(X, Y) = \Gamma_{i,0}(X) + X^{r_i} Y$. By the definition of the Puiseux transformations, the polynomial

$$G_i(X, Y) := \frac{F(\gamma_i X^{e_i}, \Gamma_i(X, Y))}{X^{N_i e_i}}$$

is such that $(0, 1) \in \mathcal{N}(G_i)$, i.e. $v_X(\partial_Y G_i(X, 0)) = 0$. Thus, $v_X(X^{r_i} F_Y(\gamma_i X^{e_i}, \Gamma_{i,0}(X))) = N_i e_i$, or:

$$N_i = \frac{r_i + v_X(F_Y(\gamma_i X^{e_i}, \Gamma_{i,0}(X)))}{e_i} = \frac{r_i}{e_i} + v_X(F_Y(X, \Gamma_{i,0}((X/\gamma_i)^{1/e_i}))) = \frac{r_i}{e_i} + v_i. \quad \square$$

Remark 5. From this result, we see that the value N_i does not depend on the algorithm. Nevertheless, the proof above rely on the algorithm RNPuiseux because it computes *precisely* the singular part of all Puiseux series thanks to the *modified* Newton polygon [28, Definition 6]. The algorithm Half-RNP introduces two differences:

- The Abhyankar's trick does not change the value of the N_i . After applying it, the next value $\frac{l}{q}$ is just the addition of the $\frac{l_i}{q_i}$ we would have found with RNPuiseux (the concerned slopes being the sequence of integer slopes that compute common terms for *all* Puiseux series, plus the next one). See Example 2 below.
- Not using the modified Newton polygon \mathcal{N}^* can only change the last value $\frac{l}{q}$, if and only if the coefficient of $X^{r/e}$ is 0. This has no impact on the proof of Lemma 6 below.

In the remaining of this paper, we will define N_i as $\frac{r_i}{e_i} + v_i$.

Example 2. Let's assume that F is an irreducible polynomial with Puiseux series $S(X) = X^{1/2} + X + X^{3/2} + X^2 + X^{9/4}$. The successive values for (l, q) are:

- $(4, 2), (2, 1), (2, 1), (2, 1)$ and $(2, 2)$ with the RNPuiseux algorithm. We thus get $N = 2 + 1 + 1 + 1 + \frac{1}{2} = \frac{11}{2}$.
- $(4, 2), (14, 2)$ with the Half-RNP algorithm (assuming high enough truncation). We thus get $N = 2 + \frac{7}{2} = \frac{11}{2}$.

Lemma 6. *Let $n_0 \in \mathbb{N}$. To compute the RPE R_i with certified precision $n_0 \geq \frac{r_i}{e_i}$, it is necessary and sufficient to run Half-RNP with truncation bound $n = n_0 + v_i$. In particular, to ensure the computation of the singular part of R_i , it is necessary and sufficient use a truncation bound $n \geq N_i$.*

Proof. First note that if we start from a polynomial H known up to X^n , then the greatest n_1 so that we can certify $H_{\Delta, \xi}$ up to X^{n_1} is precisely $n_1 = qn - l$ (see Figure 1; details can be found in [26, Proof of Lemma 2, pages 210 and 211]). This explains the truncation update of line 13 in Half-RNP .

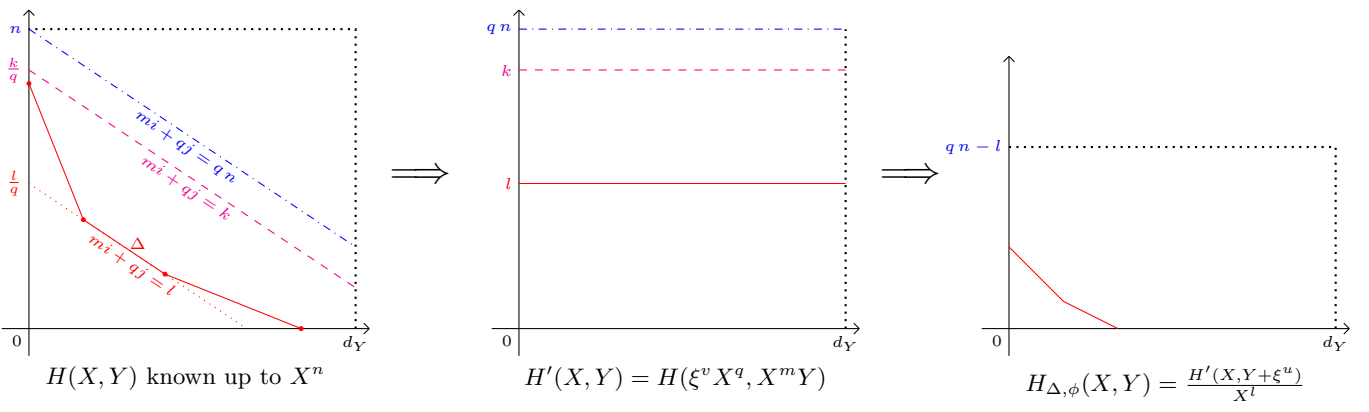


Figure 1: Change of variables for a Puiseux transform

We now distinguish two cases, according to whether the coefficient in $X^{\frac{r_i}{e_i}}$ of any Puiseux series associated to R_i is zero or not. If it is, then starting from a truncation bound $n = n' + N_i$, we get $n_1 = q n' + q N_i - l$. By definition of N_i , $q N_i - l$ is precisely the “ N_i ” of the associated RPE of $H_{\Delta, \xi} := H(\xi^v X^q, X^m (Y + \xi^u)) / X^l$. By induction, we finish at the last call of the algorithm associated to the RPE R_i with a truncation bound $n = e_i n'$. Moreover, at this point of the algorithm, we have $d_Y(H) = 1$ and $\pi = (\gamma_i X^{e_i}, \Gamma_i(X) + \alpha_i X^{r_i} Y)$. Therefore, the output of the algorithm will be R_i known with precision $n' + \frac{r_i}{e_i}$. To conclude, we just need to take $n' = n_0 - \frac{r_i}{e_i}$ thanks to Lemma 5.

Finally, if the coefficient in $X^{\frac{r_i}{e_i}}$ of any Puiseux series associated to R_i is zero, we will have $\pi = (\gamma_i X^{e_i}, \Gamma_i(X) + \alpha_i X^{\eta_i} Y)$ with $\eta_i > r_i$. If this is the case, then that means that at the previous step, we already computed some zero coefficients, thus losing the same precision $\eta_i - r_i$. This does not change the result. \square

The quantity N_i is therefore an optimal bound to compute the singular part of the RPE R_i . We now bound it.

Lemma 7. *We have $\frac{r_i}{e_i} \leq v_i$.*

Proof. This is written in the proof of [26, Proposition 5, page 204]. \square

Corollary 4. *We have $v_i \leq N_i \leq 2 v_i$.*

Proof. Straightforward consequence of Lemmas 5 and 7. \square

We finally deduce global bounds:

Proposition 6. *There are at least $\frac{d_Y}{2}$ Puiseux series $S_{i,j,k}$ such that $v_i < 2v/d_Y$ and $N_i < 4v/d_Y$.*

Proof. Let's assume that the R_i are ordered by increased v_i values, and define λ such that $\sum_{i=1}^{\lambda-1} e_i f_i < \frac{d_Y}{2} \leq \sum_{i=1}^{\lambda} e_i f_i$ (i.e. $\sum_{i=\lambda+1}^{\rho} e_i f_i \leq \frac{d_Y}{2} < \sum_{i=\lambda}^{\rho} e_i f_i$ by Proposition 1). Then we have

$$v = \sum_{i=1}^{\rho} v_i e_i f_i \geq \sum_{i=\lambda}^{\rho} v_i e_i f_i \geq v_{\lambda} \sum_{i=\lambda}^{\rho} e_i f_i > v_{\lambda} \frac{d_Y}{2},$$

the first equality being a resultant property (see e.g. [16, Exercise 6.12]). Hence, for all $i \leq \lambda$, we have $v_i \leq v_{\lambda} < 2v/d_Y$ with moreover $N_i < 4v/d_Y$ by Corollary 4. The claim follows. \square

3.4 Complexity results and proof of Theorem 4.

Proposition 7. *Not taking into account the cost of univariate factorisations, running $\text{Half-RNP}(F, Z, n, (X, Y))$ with $F \in \mathbb{K}[X, Y]$ takes less than $\mathcal{O}(M(n d_Y^2) \log(d_Y))$ arithmetic operations over \mathbb{K} .*

Proof. First note that recursive call of line 4 is never done twice in a row thanks to Lemma 3. Representing the set of recursive calls as a tree \mathcal{T} , we can group Abhyankar’s recursive call with the next Puiseux transform as a single node.

Let’s consider a function call to $\text{Half-RNP}(H, P, n_H, \pi)$ and denote $d_P = \deg_Z(P)$. We distinguish two kind of lines:

(Type 1) By Lemma 4, the number of arithmetic operations for line 4 is bounded by $\mathcal{O}(M(n d_Y))$ using the bound $n d_Y \geq n_H d_Y(H) d_P$. Also, from Remark 3 (and respectively Lemmas 2 and 3), Lines 3 and 12 are less than $\mathcal{O}(M(n d_Y))$.

(Type 2) By respectively Lemma 2 and Proposition 4, lines 13 and 14 are less than $\mathcal{O}(M(q_\Delta d_\phi n d_Y))$. Thanks to Proposition 5, line 14 fits in the same complexity bound while line 12 costs $\mathcal{O}((d_P d_\phi)^{\frac{\omega+1}{2}})$.

From Lemma 3, when $q = d_\phi = 1$, we must have a branch separation for the corresponding node of \mathcal{T} . Therefore, this happens at most $\rho - 1$ times (more precisely, the sum of the nodes of \mathcal{T} of couples (Δ, ϕ) with $q = d_\phi = 1$ is bounded by $\rho - 1$). This means that the sum of the costs for these cases is less than $\mathcal{O}(\rho M(n d_Y))$.

To conclude the proof, we still have to deal with all the cases where $q > 1$ or $d_\phi > 1$. In such a case, Type 2 lines are the costly ones. Moreover, we can bound q by e_i and $d_P d_\phi$ by f_i for any RPE R_i issued from (Δ, ϕ) . But for each RPE R_i , such situation cannot happen most than $\log(e_i f_i) \leq \log(d_Y)$ times (before and/or after separation of this branch with other ones). From Definition 10, that means we can bound the total cost for all these cases by $\mathcal{O}((M(\sum_{i=1}^{\rho} e_i f_i n d_Y) + \sum_{i=1}^{\rho} f_i^{\frac{\omega+1}{2}}) \log(d_Y)) \subset \mathcal{O}(M(n d_Y^2) \log(d_Y))$. \square

Proof of Theorem 5. As far as correctness is concerned, we only have to take care of truncations and the precision of the output. Other points have been considered in previous papers of the first author [24, 26, 28] (note also [13, Section 4.1] concerning the construction of the output). From Proposition 6 and Lemma 6, we know that we will get at least half of the Puiseux series with precision $4v/d_Y$ or greater and $v_i < 2v/d_Y$ by a function call $\text{Half-RNP}(F, Z, 6v/d_Y, (X, Y))$. Finally, the complexity follows from Proposition 7. \square

4 A dichotomic Newton-Puiseux algorithm.

We now provide a dichotomic strategy to compute all Puiseux series.

Theorem 6. *Let $F \in \mathbb{K}[X][Y]$ a primitive and separable polynomial. Assume that $p = 0$ or $p > d_Y$. Not taking into account the cost of univariate factorisations, there exists an algorithm that computes the singular part of all rational Puiseux expansions above $X = 0$ in less than $\mathcal{O}(M(d_Y v) \log(d_Y v) + M(d_Y) \log^2(d_Y))$ arithmetic operations.*

Assuming that F is monic, our strategy can be summarized as follows:

1. We run $\text{Half-RNP}(F, Z, 6v/d_Y, (X, Y))$. If this provides us all RPEs of F , we are done. If not, from Section 3, we get at least half of the Puiseux series of F , with precision $4v/d_Y$ or more and with $v_i < 2v/d_Y$.
2. From these Puiseux series, we construct the associated irreducible factors and their product G with precision $4v/d_Y$; this is detailed in Section 4.1. Note that $d_Y(G) \geq d_Y/2$ thanks to Theorem 5.
3. We compute its cofactor H by euclidean division modulo X^{4v/d_Y+1} .
4. We compute a Bezout relation $UG + VH = X^k \pmod{X^{k+1}}$ using [19, Algorithm 1]. We additionally prove that there exists such k with $k \leq 2v/d_Y$. This is Idea 4, detailed in Section 4.2.
5. From this relation and the factorisation $F = GH \pmod{X^{4v/d_Y+1}}$, we will adapt the Hensel lemma to get a factorisation $F = GH \pmod{X^{v+1}}$ in quasi-linear time. This is Idea 3, detailed in Section 4.3.
6. Finally, we apply our main algorithm recursively on H ; as the degree in Y is at least divided by two each time, this is done at most $\log(d_Y)$ times, for a total cost only multiplied by 2. This is Idea 5, detailed in Section 4.4.

Note that F is not assumed monic in Theorem 6. If it is not, we will first use Hensel lifting to compute the factor F_∞ corresponding to RPEs centered at $(0, \infty)$ up to precision X^v . Then, we will compute the RPEs of F_∞ as “inverse” of the RPEs of its reciprocal polynomial (which is monic by construction). Details are provided in Section 4.5.

4.1 Computing the norm of a RPE.

Lemma 8. *Let R_1, \dots, R_λ be a set of \mathbb{K} -RPEs not centered at $(0, \infty)$. For each R_i , we denote $(S_{ijk})_{jk}$ its associated Puiseux series. Let*

$$\nu = \max_{1 \leq i \leq \lambda} \sum_{\substack{(i', j', k') \\ \neq (i, j, k)}} v_X(S_{ijk}(X) - S_{i'j'k'}(X))$$

and assume that the R_i are known with precision $n \geq \nu$. Then there exists an algorithm `NormRPE` that computes a monic polynomial $G \in \mathbb{K}[X, Y]$ with $d_Y(G) = \sum_{i=1}^{\lambda} e_i f_i$, $d_X(G) = n + \nu$, and such that the RPE of G with precision n are precisely the R_i . It takes less than $\mathcal{O}(M(n d_Y(G)^2) \log(n d_Y(G))) \subset \mathcal{O}(n d_Y(G)^2)$ arithmetic operations over \mathbb{K} .

Proof. Denote $P_i \in \mathbb{K}[Z]$ so that $R_i = (\gamma_i(Z) T^{e_i}, \Gamma_i(Z, T))$ is defined over \mathbb{K}_{P_i} . For $1 \leq i \leq \lambda$, compute

$$A_i = \prod_{j=0}^{e_i-1} \left(Y - \Gamma_i \left(Z, \zeta_{e_i}^j \left(\frac{X}{\gamma_i} \right)^{\frac{1}{e_i}} \right) \right) \pmod{(X^{n+\nu+1}, P_i(Z))}$$

As $n \geq \nu$, it takes $\mathcal{O}(M(e_i^2 n f_i) \log(e_i))$ operations in \mathbb{K} using a subproduct tree. Then, we compute $G_i = \text{Res}_Z(A_i, P_i) \pmod{X^{n+\nu+1}}$ in $\mathcal{O}(f_i M(n e_i f_i) \log(n e_i f_i))$ (just adapt [16, Corollary 11.21, page 332] to a polynomial with three variables). Summing over i these two operations, this fits into our bound. Finally, we compute the product of the G_i modulo $X^{n+\nu+1}$ in less than $\mathcal{O}(M(n d_Y(G)) \log(d_Y(G)))$ using a subproduct tree. \square

4.2 Lifting order.

Our algorithm requires to lift some analytic factors G, H of F which are not coprime modulo (X) . To this aim, we will generalise the classical Hensel lifting. The first step is to compute a generalized Bezout relation $UG + VH = X^\eta$ for some η as small as possible.

Definition 11. Let $G, H \in \mathbb{K}[[X]][Y]$ coprime. The *lifting order* of G and H is defined as:

$$\kappa(G, H) := \inf \{ \eta \in \mathbb{N}, X^\eta \in (G, H) \}.$$

The following proposition gives an upper bound for the lifting order sufficient for our purpose.

Proposition 8. *Suppose H monic. We have $\kappa(G, H) \leq \max_{H(S)=0} v_X(F_Y(S))$, where $F = G \cdot H$.*

Proof. Let $UG + VH = X^\kappa$ in $\mathbb{K}[[X]][Y]$, with $\kappa = \kappa(G, H)$ minimal. Up to perform euclidean division of U by H , we may assume that $d_Y(U) < d_Y(H) = d$. Moreover, minimality of κ and monicity of H impose $v_X(U) = 0$. Denoting S_1, \dots, S_d the Puiseux series of H , we have $U(S_i)G(S_i) = X^\kappa$ for $1 \leq i \leq d$. Using interpolation, we get

$$U = \sum_{i=1}^d \frac{X^\kappa}{G(S_i)H_Y(S_i)} \prod_{j \neq i} (Y - S_j) = \sum_{i=1}^d \frac{X^\kappa}{F_Y(S_i)} \prod_{j \neq i} (Y - S_j).$$

Using $v_X(U) = 0$ and $v_X(S_j) \geq 0$ (H is monic), we deduce that $\kappa \leq \max_{1 \leq i \leq d} v_X(F_Y(S_i))$, as required. \square

Corollary 5. *Assume that $F \in \mathbb{K}[[X]][Y]$ is a non irreducible monic polynomial. Then there exists a factorisation $F = GH$ in $\mathbb{K}[[X]][Y]$ such that $\kappa(G, H) \leq 2v/d_Y$.*

Proof. From Proposition 6, there exist $\lambda \geq 1$ RPE R_1, \dots, R_λ of F such that $v_i < 2v/d_Y$ for all $i \leq \lambda$. If we consider $H = \prod_{i=1}^{\lambda} F_i$ and $G = \prod_{i=\lambda+1}^{\rho} F_i$, where F_i is the analytic factor associated to R_i (see Section 2.1), then we are done thanks to Proposition 8. \square

Finally, note that the relation $UG + VH = X^\eta \pmod{X^{\eta+1}}$ can be computed in $\mathcal{O}(M(d_Y \eta) \log(\eta) + M(d_Y) \eta \log(d_Y))$ from [19, Corollary 1], that is less than $\mathcal{O}(M(v) \log(v))$ if we consider the factorisation of Corollary 5.

4.3 Adaptation of Hensel's lemma to our context.

We now generalise the classical Hensel lemma, starting from a relation $F = GH$ known modulo $X^{2\eta+1}$ with a Bezout relation $UG + VH = X^\eta \pmod{X^{\eta+1}}$. We refer for instance to [16, section 15.4, page 444] for the classical version when G and H are coprime modulo (X) . Note that QuoRem in here is the classical euclidean division algorithm.

Algorithm: HenselStep(F, G, H, U, V, n_0, η)

Input: $F, G, H, U, V \in \mathbb{K}[Y][X]$, $n_0, \eta \in \mathbb{N}$ s.t. $n_0 > 2\eta$, $F = G \cdot H \pmod{X^{n_0}}$, $U \cdot G + V \cdot H = X^\eta \pmod{X^{n_0-\eta}}$, $d_Y(U) < d_Y(H)$, $d_Y(V) < d_Y(G)$ and H monic.

Output: $\tilde{G}, \tilde{H}, \tilde{U}, \tilde{V} \in \mathbb{K}[Y][X]$ s.t. $\tilde{G} = G \pmod{X^{n_0-\eta}}$, $\tilde{H} = H \pmod{X^{n_0-\eta}}$, $F = \tilde{G} \cdot \tilde{H} \pmod{X^{2(n_0-\eta)}}$, $\tilde{U} \cdot \tilde{G} + \tilde{V} \cdot \tilde{H} = X^\eta \pmod{X^{2n_0-3\eta}}$, $d_Y(\tilde{V}) < d_Y(\tilde{G})$, $d_Y(\tilde{U}) < d_Y(\tilde{H}) = d_Y(H)$ and \tilde{H} monic.

- 1 $\alpha \leftarrow X^{-\eta}(F - G \cdot H) \pmod{X^{2(n_0-\eta)}}$;
- 2 $Q, R \leftarrow \text{QuoRem}_Y(U \cdot \alpha, H) \pmod{X^{2(n_0-\eta)}}$;
- 3 $\tilde{G} \leftarrow G + \alpha \cdot V + Q \cdot G \pmod{X^{2(n_0-\eta)}}$;
- 4 $\tilde{H} \leftarrow H + R \pmod{X^{2(n_0-\eta)}}$;
- 5 $\beta \leftarrow X^{-\eta}(U \cdot \tilde{G} + V \cdot \tilde{H}) - 1 \pmod{X^{2n_0-3\eta}}$;
- 6 $S, T \leftarrow \text{QuoRem}_Y(U \cdot \beta, \tilde{H}) \pmod{X^{2(n_0-\eta)}}$;
- 7 $\tilde{U} \leftarrow U - T \pmod{X^{2n_0-3\eta}}$;
- 8 $\tilde{V} \leftarrow V - \beta \cdot V - S \cdot \tilde{G} \pmod{X^{2n_0-3\eta}}$;
- 9 **return** $\tilde{H}, \tilde{G}, \tilde{U}, \tilde{V}$

Lemma 9. *Algorithm HenselStep is correct; it takes less than $\mathcal{O}(M(n_0 d_Y))$ operations in \mathbb{K} .*

Proof. From $\alpha \equiv 0 \pmod{X^{n_0-\eta}}$ (thus $Q \equiv 0 \pmod{X^{n_0-\eta}}$ and $R \equiv 0 \pmod{X^{n_0-\eta}}$ from [16, Lemma 15.9, (ii), page 445]) and $U \cdot G + V \cdot H - X^\eta \equiv 0 \pmod{X^{n_0-\eta}}$, we have $\tilde{G} \equiv G \pmod{X^{n_0-\eta}}$, $\tilde{H} \equiv H \pmod{X^{n_0-\eta}}$ and

$$\begin{aligned} F - \tilde{G} \cdot \tilde{H} &\equiv F - (G + \alpha \cdot V + Q \cdot G) \cdot (H + \alpha \cdot U - Q \cdot H) \\ &\equiv \alpha(X^\eta - V \cdot H - U \cdot G) - \alpha^2 \cdot U \cdot V - Q \cdot \alpha(U \cdot G - V \cdot H) + Q^2 \cdot G \cdot H \equiv 0 \pmod{X^{2(n_0-\eta)}}. \end{aligned}$$

From $\beta \equiv 0 \pmod{X^{n_0-2\eta}}$ and $U \cdot \tilde{G} + V \cdot \tilde{H} - X^\eta \equiv 0 \pmod{X^{n_0-\eta}}$, we have:

$$\begin{aligned} \tilde{U} \cdot \tilde{G} + \tilde{V} \cdot \tilde{H} - X^\eta &\equiv (U - U \cdot \beta + S \cdot \tilde{H}) \cdot \tilde{G} + (V - \beta \cdot V - S \cdot \tilde{G}) \cdot \tilde{H} - X^\eta \\ &\equiv U \cdot \tilde{G} + V \cdot \tilde{H} - X^\eta - \beta \cdot (U \cdot \tilde{G} + V \cdot \tilde{H}) \\ &\equiv \beta \cdot (X^\eta - U \cdot \tilde{G} - V \cdot \tilde{H}) \equiv 0 \pmod{X^{2n_0-3\eta}}. \end{aligned}$$

Conditions on the degrees in Y for \tilde{H} and \tilde{U} are obvious (thus is the monicity of \tilde{H}). The complexity result is similar to [16, Theorem 9.6, page 261]. \square

Lemma 10. *Given F, G, H as in the input of algorithm HenselStep with $n_0 = 2\eta + 1$, there exists an algorithm that computes polynomials as in the output of HenselStep for any precision $n \in \mathbb{N}$, additionally satisfying:*

- $\tilde{G} = G \pmod{X^{\eta+1}}$, $\tilde{H} = H \pmod{X^{\eta+1}}$ and $F = \tilde{G} \cdot \tilde{H} \pmod{X^{n+2\eta}}$;
- if $\exists G^*, H^* \in \mathbb{K}[Y][X]$ s.t. $F = G^* \cdot H^* \pmod{X^{n+2\eta}}$, then $\tilde{G} = G^* \pmod{X^n}$ and $\tilde{H} = H^* \pmod{X^n}$.

We call Hensel such an algorithm. It takes less than $\mathcal{O}(M(n d_Y) + M(\eta d_Y) \log(\eta d_Y))$ operations in \mathbb{K} .

Proof. The algorithm runs as follows:

1. Compute $U, V \in \mathbb{K}[X, Y]$ such that $U \cdot G + V \cdot H = X^\eta \pmod{X^{\eta+1}}$ [19, Algorithm 1].
2. Double the value $n_0 - 2\eta$ at each call of HenselStep, until we get $n_0 - 2\eta \geq n + \eta$.

Correctness and complexity are a straightforward consequence of Lemma 9 (using [19, Corollary 1] for the computation of U and V). Finally, uniqueness of the result is an adaptation of [16, Theorem 15.14, page 448] (this works because we take a precision satisfying $n_0 - 2\eta \geq n + \eta$). \square

Remark 6. Note that if $G(0, Y)$ and $H(0, Y)$ are coprime, then $\eta = 0$ and this result is the classical Hensel lemma.

4.4 The divide and conquer algorithm for monic polynomials.

We now provide our divide and conquer algorithm. Algorithm Quo therein outputs the quotient of the euclidean division in $\mathbb{K}[[X]][Y]$ modulo a power of X .

Algorithm: MonicRNP(F, n)

Input: $F \in \mathbb{K}[X, Y]$, separable and monic in Y ; $n \in \mathbb{N}$.

Output: If n is big enough, the singular part (at least) of all the RPEs of F above $X = 0$.

```

1 if  $d_Y < 6$  then return Half-RNP( $F, Z, n, (X, Y)$ ) else  $\eta \leftarrow 6n/d_Y$ ;
2  $\mathcal{R} \leftarrow$  Half-RNP( $F, Z, \eta, (X, Y)$ );
3 Keep in  $\mathcal{R}$  the RPEs with  $v_i < \eta/3$ ; // these RPEs are known with precision  $\geq 2\eta/3$ 
4 if  $\sum_{i=1}^{\lambda} e_i f_i = d_Y$  then return  $\mathcal{R}$ ; //  $\mathcal{R} = \{R_1, \dots, R_\lambda\}$ ;  $e_i$  and  $f_i$  associated to  $R_i$ 
5  $G \leftarrow$  NormRPE( $\mathcal{R}, 2\eta/3$ );
6  $H \leftarrow$  Quo( $F, G, 2\eta/3$ );
7  $G, H \leftarrow$  Hensel( $F, G, H, n$ );
8 return  $\mathcal{R} \cup$  MonicRNP( $H, n$ );

```

Proposition 9. *If F is monic, a function call MonicRNP(F, n) with $n \geq v$ computes all singular parts (at least) of rational Puiseux expansions of F above 0. Without taking into account the cost of univariate factorisations, it takes less than $\mathcal{O}(M(d_Y n) \log(d_Y n))$ arithmetic operations.*

Proof. If $d_Y < 6$, Line 1 is $\mathcal{O}(M(n))$ by Proposition 7, and precision $n \geq v$ is enough to get the singular part of the Puiseux series. From Proposition 7, line 2 takes $\mathcal{O}(M(n d_Y) \log(d_Y))$. Moreover, from Lemma 6, the R_i , $1 \leq i \leq \lambda$, are known with precision $\eta - v_i \geq 2\eta/3$; and we have $d_Y(G) \geq d_Y/2 \geq d_Y(H)$ from Proposition 6. Then, lines 5, 6 and 7 take respectively less than $\mathcal{O}(M(n d_Y) \log(n d_Y))$, $\mathcal{O}(M(n d_Y))$ and $\mathcal{O}(M(n d_Y) + M(v) \log(v))$ by respectively Lemma 8, division via Newton iteration [16, Theorem 9.4, page 260] and Lemma 10 (input of the Hensel algorithm is correct since $\kappa(G, H)$ is less than $\eta/3$ by Proposition 8 and we have $F = G \cdot H \pmod{X^{2\eta/3+1}}$). This fits into our bound. Finally, as $d_Y(H) \leq d_Y/2$, complexity follows from Lemma 1. \square

4.5 Dealing with the non monic case: proof of Theorem 6.

Proposition 10. *There exists an algorithm Monic that given $n \in \mathbb{N}$ and $F \in \mathbb{K}[X][Y]$ primitive, returns two polynomials $F_0, F_\infty \in \mathbb{K}[X][Y]$ such that $F = u F_0 F_\infty \pmod{X^n}$, with F_0 monic in Y , $F_\infty(0, Y) = 1$, and $u \in \mathbb{K}[X]$, $u(0) \neq 0$ with at most $\mathcal{O}(M(n d_Y))$ operations over \mathbb{K} .*

Proof. This is [20, Algorithm Q, page 33] (see the proof of Proposition 4). \square

We can now give our main algorithm RNP. It computes the singular part of all RPEs of F above $X = 0$, including those centered at $(0, \infty)$.

Algorithm: RNP(F, n)

Input: $F \in \mathbb{K}[X, Y]$, separable in Y and $n \in \mathbb{N}$.

Output: If n is big enough, the singular part (at least) of all the RPEs of F above $X = 0$

```

1  $(F_0, F_\infty) \leftarrow$  Monic( $F, n$ );
2  $F'_\infty \leftarrow Y^{d_Y(F_\infty)} F_\infty(X, 1/Y)$ ;
3  $\mathcal{R}_\infty \leftarrow$  MonicRNP( $F'_\infty, n$ );
4 Inverse the second element of each  $R \in \mathcal{R}_\infty$ ; // see the proof of Proposition 11 for details
5 return MonicRNP( $F_0, n$ )  $\cup$   $\mathcal{R}_\infty$ ;

```

The proof of Theorem 6 follows immediately from the following proposition:

Proposition 11. Assume that $p = 0$ or $p > d_Y$. Not taking into account the cost of univariate factorisations, a function call $\text{RNP}(F, v)$ returns the expected output with at most $\mathcal{O}(M(d_Y v) \log(d_Y v))$ arithmetic operations.

There is one delicate point in the proof of Proposition 11: we need to invert the RPEs of F'_∞ and it is not clear that the truncation bound $n = v$ is sufficient for recovering in such a way the singular part of the RPEs of F_∞ (see also Remark 7 below). We will need the following two results:

Lemma 11. Given two distinct Puiseux series S and S' , we have $v_X\left(\frac{1}{S} - \frac{1}{S'}\right) = v_X(S - S') - v_X(S) - v_X(S')$.

Proof. If $v_X(S) \neq v_X(S')$, one can assume up to renaming the series that $v_X(S) < v_X(S')$. Then $v_X(S - S') = v_X(S)$ and $v_X\left(\frac{1}{S} - \frac{1}{S'}\right) = v_X(S')$. If $v_X(S) = v_X(S') = \alpha$, then $\frac{X^\alpha}{S} \bmod X^n$ is uniquely determined from $X^{-\alpha} S \bmod X^n$ (same for S'). Denoting $s = v_X(S - S') - \alpha$, then we have $X^{-\alpha} S = X^{-\alpha} S' \bmod X^s$, thus $\frac{X^\alpha}{S} = \frac{X^\alpha}{S'} \bmod X^s$ and $\alpha + v_X\left(\frac{1}{S} - \frac{1}{S'}\right) \geq s$. Similarly, denoting $s_0 = \alpha + v_X\left(\frac{1}{S} - \frac{1}{S'}\right)$, we get $v_X(S - S') - \alpha \geq s_0$, concluding the proof. \square

Proposition 12. Let $F_\infty \in \mathbb{K}[X, Y]$ satisfying $F_\infty(0, Y) = 1$ and denote F'_∞ its reciprocal polynomial according to the variable Y . For each RPE $R_i = (\lambda_i X^{e_i}, \Gamma_i)$ of F_∞ , denote $s_i := v_X(\Gamma_i)$ (so $s_i < 0$), r_i its regularity index and R'_i the associated RPE of F'_∞ . The function call $\text{MonicRNP}(F'_\infty, v_{F_\infty})$ computes each RPE R'_i with precision at least $\frac{r_i - 2s_i}{e_i}$.

Proof. Denote $d := d_Y(F_\infty)$ and let v stands for the X -valuation of the leading coefficient of F_∞ with respect to Y . Let S_1, \dots, S_d the Puiseux series of F_∞ and S_{k_i} one of the Puiseux series associated to the RPE R_i of F_∞ we are considering. Note that $\frac{s_i}{e_i} = v_X(S_{k_i})$. By definition of the regularity index, we have $v_X(S_{k_i} - S_j) \leq \frac{r_i}{e_i}$ for all $j \neq k_i$. Let i_0 satisfying $v_X(S_{k_i} - S_{i_0}) = \max_{j \neq k_i} v_X(S_{k_i} - S_j)$ (several values of i_0 are possible). We distinguish three cases:

1. $v_X(S_{k_i}) = v_X(S_{i_0})$; then we either have $v_X(S_{k_i} - S_{i_0}) = \frac{r_i}{e_i}$, or $e_{i_0} = q e_i$ with q an integer > 1 . In the latter case, there exist q conjugates Puiseux series $(S_{i_0}^{[j]})_{0 \leq j < q}$ of S_{i_0} such that $v_X(S_{k_i} - S_{i_0}) = v_X(S_{k_i} - S_{i_0}^{[j]})$. Therefore, $\sum_{j=0}^{q-1} v_X(S_{k_i} - S_{i_0}^{[j]}) \geq \frac{r_i}{e_i}$; see [26, Case 3 in Proof of Proposition 5, pages 204 and 205] for details.
2. $v_X(S_{k_i}) > v_X(S_{i_0})$. Then $r_i = s_i$ and $v_X(S_{k_i}) > v_X(S_j)$ for any $j \neq k_i$ by definition of i_0 . In particular, this induces $\frac{v}{d} \geq -\frac{s_i}{e_i}$ since $v = \sum_{k=1}^d -\frac{s_k}{e_k}$ from [26, Lemma 1, page 198].
3. $v_X(S_{k_i}) < v_X(S_{i_0})$. Then $v_X(S_{k_i} - S_{i_0}) = s_i = r_i$. We can also assume that $v_X(S_j) \neq v_X(S_{k_i})$ for all $j \neq k_i$: if $v_X(S_j) = v_X(S_{k_i})$, then $v_X(S_{k_i} - S_j) = v_X(S_{k_i} - S_{i_0})$ and one could use $i_0 = j$ and deal with it as Case 1.

We can now prove Proposition 12. First, when considering Case 1, knowing $\frac{1}{S_{k_i}}$ with precision $v'_i := v_X\left(\partial_Y F'_\infty\left(\frac{1}{S_{k_i}}\right)\right)$ is sufficient: from Lemma 11, we have $v'_i = \sum_{j \neq i} v_X\left(\frac{1}{S_{k_i}} - \frac{1}{S_j}\right) = \sum_{j \neq i} v_X(S_{k_i} - S_j) - v_X(S_{k_i}) - v_X(S_j)$. As either $v_X(S_{k_i} - S_{i_0}) - v_X(S_{k_i}) - v_X(S_{i_0}) = \frac{r_i - 2s_i}{e_i}$ or $\sum_{j=0}^{q-1} v_X(S_{k_i} - S_{i_0}^{[j]}) - v_X(S_{k_i}) - v_X(S_{i_0}^{[j]}) \geq \frac{r_i - 2s_i}{e_i}$, we are done.

Then, concerning Case 2, from Proposition 9 and Lemma 6, we know the RPE R_i with precision at least $v_i + \frac{v}{d}$, that is at least $\frac{v}{d}$ thus $-\frac{s_i}{e_i}$ from the discussion above. As $r_i = s_i$, this is at least $\frac{r_i - 2s_i}{e_i}$.

Finally, Case 3 requires more attention. Let's first assume that $v_X(S_{k_i}) > v_X(S_j)$ for some $j \neq i$; then $v_X(S_{k_i} - S_j) - v_X(S_{k_i}) - v_X(S_j) = v_X(S_{k_i}) = -\frac{s_i}{e_i}$, and we are done since $r_i = s_i$. If not, then we have $v_X(S_{k_i}) < v_X(S_j)$ for all j . This means that $e_i = f_i = 1$, and that $\mathcal{N}(F'_\infty)$ has an edge $[(0, v), (1, v - s_i)]$, which is associated to R'_i . It is enough to prove that the truncation bound used when dealing with this Puiseux series is at least v . As long as this is not the case, this edge is not considered from the definition of $\mathcal{N}_n(H)$; also, at each recursive call of MonicRNP (Line 8), the value of the truncation bound η increases (since the degree in Y is at least divided by 2). In the worst case, we will end with a polynomial of degree 1, using a truncation bound $\eta = v_{F_\infty} \geq v$. This concludes. \square

Proof of Proposition 11. Let us show that the truncation bound for F'_∞ is sufficient for recovering the singular part of the Puiseux series of F_∞ . First note that the inversion of the second element is done as follows: consider $R'_i(T) = (\gamma_i T^{e_i}, \Gamma'_i(T) = \sum_{k=0}^{\tau_i} \alpha'_{i,k} T^k)$ and denote $s_i = -v_T(\Gamma'_i(T)) < 0$; we compute the inverse of $T^{s_i} \Gamma'_i(T)$ (that has a non zero constant coefficient) via quadratic Newton iteration [16, Algorithm 9.3, page 259]; it takes less than $\mathcal{O}(M(\tau_i + s_i))$ arithmetic operations [16, Theorem 9.4, page 260]. In order to get the singular part of the corresponding RPE R_i of F_∞ , we need to know R_i with precision $\frac{r_i}{e_i}$, i.e. to know at least $r_i - s_i + 1$ terms. That means that we need to know R'_i with precision $r_i - 2s_i$ or more. This holds thanks to Proposition 12. Correctness and complexity of Algorithm RNP then follows straightforwardly from Propositions 9 and 10. \square

Remark 7. Note that precision $v_X(\text{Disc}_Y F)$ is not always enough to get the singular part of the Puiseux series centered at $(0, \infty)$, as shows the following example. We consider $F(X, Y) = 1 + X Y^{d-1} + X^{d+1} Y^d$. The singular parts of its RPEs are $(T, \frac{-1}{T^d})$ and $(-1 T^{d-1}, \frac{1}{T})$. Its reciprocal polynomial is $F' = Y^d + X Y + X^{d+1}$, with RPE's singular parts $(T, 0)$ and $((-1)^d T^{d-1}, -T)$. Here we have $v_X(\text{Disc}_Y F) = d$, and $[F']^d = Y^d + X Y$. The singular parts of $[F']^d$ are indeed the same than the one of F' , but we cannot recover the RPE $(T, \frac{-1}{T^d})$ of F from the RPE $(T, 0)$ of $[F']^d$. Nevertheless, the precision $v = v_X(\text{lc}_Y(F)) + v_X(\text{Disc}_Y F) = 2d + 1$ is sufficient.

Proof of Theorem 6. The algorithm mentionned in this theorem is RNP described above, run with parameter (F, v) . We can compute the parameter v with at most $\mathcal{O}(M(d_Y v) \log(d_Y v))$ operations from [19, Corollary 1]. \square

Remark 8. Another way to approach the non monic case is the one used in [26]. The idea is to use algorithms `MonicRNP` and `Half-RNP3` even when F is not monic. This would change nothing as far as these algorithms are concerned, but the proof concerning truncation bounds must be adapted:

1. define $s_i := \min(0, v_X(S_i))$, $N'_i := N_i - \frac{s_i}{e_i} d_Y$ and $v'_i := v_i - \frac{s_i}{e_i} d_Y$;
2. prove that $N'_i = \frac{r_i}{e_i} + v'_i$ (use [26, Figure 3, page 211] to deal with the possible positive slopes of the initial call);
3. replace v_i by v'_i and N_i by N'_i in the remaining results of Section 3.3; proofs use some intermediate results of [26] (in particular, to prove $\frac{r_i}{e_i} \leq v'_i$, we need to use some formulas in the proof of [26, Proposition 5, page 204]).

We chose to consider the monic case separately, since it makes one of the main technical results of this paper (namely tight truncation bounds) less difficult to apprehend, thus the paper more progressive to read.

5 Avoiding univariate factorisation.

As pointed out in the conclusion of [28], even over finite fields (when considering *all* critical points), the cost of univariate factorisation cannot be bounded so that it fits into our objectives. Up to now, we proved Theorem 1 up to the cost of univariate factorisations. To conclude the proof, one would additionally need to prove that the cost of all univariate factorisations computed when calling Algorithm `Half-RNP` is in $\mathcal{O}(v d_Y)$. As v can be small, we would need a univariate factorisation algorithm for a polynomial in $\mathbb{K}[Y]$ of degree at most d with complexity $\mathcal{O}(d)$. Unfortunately, this does not exist. We will solve this point via Idea 6; relying on the “dynamic evaluation” technique [11, 10] (also named “D5 principle”) of Della Dora, Dicrescenzo and Duval. This provides a way to compute with algebraic numbers, while avoiding factorisation (replacing it by *square-free* factorisation). In this context, we will consider polynomials with coefficients belonging to a direct product of field extensions of \mathbb{K} ; more precisely to a zero-dimensional *non integral* \mathbb{K} -algebra $\mathbb{K}_I = \mathbb{K}[\underline{Z}]/I$, where I is defined as a triangular set in $\mathbb{K}[\underline{Z}] := \mathbb{K}[Z_1, \dots, Z_s]$. As a consequence, zero divisors might appear, causing triangular decomposition and splittings (see Section 5.1 for details). Three main subroutines of the `Half-RNP` algorithm can lead to a decomposition of the coefficient ring:

- (i) computation of Newton polygons,
- (ii) square-free factorisations of characteristic polynomials,
- (iii) computation of primitive elements.

There are two other points that we need to take care of for our main program:

- (iv) subroutine `Hensel`, via the initial use of [19, Algorithm 1];
- (v) the initial factorisation of algorithm RNP (when computing Puiseux series above *all* critical points).

To simplify the comprehension of this section, we will not mention logarithm factors in our complexity results, using only the \mathcal{O} notation. This section is divided as follows:

1. We start by recalling a few definitions on triangular sets and in particular our notion of *D5 rational Puiseux expansions* in Section 5.1.
2. The key point of this section is to deal with these splitting with almost linear algorithms; to do so, we mainly rely on [10]. We briefly review in Section 5.2 their results; additionally, we introduce a few algorithms needed in our context. In particular, this section details points (iii) and (iv) above.

3. Points (i) and (ii) above are grouped in a unique procedure `Polygon-Data`, detailed in Section 5.3.
4. We provide the D5 versions of algorithms `Half-RNP`, `MonicRNP` and `RNP` in Section 5.4.
5. Finally, we prove Theorem 1 in Section 5.5.

5.1 Triangular sets and dynamic evaluation.

Definition 12. A (monic, autoreduced) *triangular set* of $\mathbb{K}[Z_1, \dots, Z_s]$ is a set of polynomials P_1, \dots, P_s such that:

- $P_i \in \mathbb{K}[Z_1, \dots, Z_i]$ is monic in Z_i ,
- P_i is reduced modulo (P_1, \dots, P_{i-1}) ,
- the ideal (P_1, \dots, P_s) of $\mathbb{K}[Z]$ is radical.

We abusively call an ideal $I \subset \mathbb{K}[Z]$ a triangular set if it can be generated by a triangular set (P_1, \dots, P_s) . We denote by \mathbb{K}_I the quotient ring $\mathbb{K}[Z]/(I)$.

Note that this defines a zero-dimensional lexicographic Gröbner basis for the order $Z_1 < \dots < Z_s$ with a triangular structure. Such a product of fields contains zero divisor:

Definition 13. We say that a non-zero element $\alpha \in \mathbb{K}_I$ is *regular* if it is not a zero divisor. We say that a polynomial or a parametrisation defined over \mathbb{K}_I is regular if all its non zero coefficients are regular.

Given a zero divisor α of \mathbb{K}_I , one can divide I as $I = I_0 \cap I_1$ with $I_0 + I_1 = (1)$, $\alpha \bmod I_0 = 0$ and $\alpha \bmod I_1$ is invertible. Moreover, both ideals I_0 and I_1 can be represented by triangular sets of $\mathbb{K}[Z]$.

Triangular decomposition.

Definition 14. A *triangular decomposition* of an ideal I is $I = I_1 \cap \dots \cap I_k$ such that $I_i + I_j = (1)$ for any $1 \leq i \neq j \leq k$, and such that every I_i can be represented by a triangular set.

Thanks to the Chinese remainder theorem, the \mathbb{K} -algebra \mathbb{K}_I is isomorphic to $\mathbb{K}_{I_1} \oplus \dots \oplus \mathbb{K}_{I_k}$ for any triangular decomposition of I . We extend this isomorphism coefficient wise for any polynomial or series defined above \mathbb{K}_I .

Definition 15. Consider any polynomial or series defined above \mathbb{K}_I . We define its *splitting* according to a triangular decomposition $I = I_1 \cap \dots \cap I_k$ the application of the above isomorphism coefficient-wise.

A key point (as far complexity is concerned) is the concept of *non critical* triangular decompositions. We recall [10, Definitions 1.5 and 1.6]:

Definition 16. Two polynomials $a, b \in \mathbb{K}_I[X]$ are said *coprime* if the ideal $(a, b) \subset \mathbb{K}_I[X]$ is equal to (1) .

Definition 17. Let (P_1, \dots, P_s) and (P'_1, \dots, P'_s) be two distinct triangular sets. We define the *level* l of this two triangular sets to be the least integer such that $P_l \neq P'_l$. We say that these triangular sets are *critical* if P_l and P'_l are not coprime in $\mathbb{K}[Z_1, \dots, Z_{l-1}]/(P_1, \dots, P_{l-1})$. A triangular decomposition $I = I_1 \cap \dots \cap I_k$ is said *non critical* if it has no critical pairs ; otherwise, it is said *critical*.

D5 rational Puiseux expansions. We conclude this section by defining systems of D5-RPEs over fields and product of fields. Roughly speaking, a system of D5-RPE over a field \mathbb{K} is a system of RPEs over \mathbb{K} grouped together with respect to some square-free factorisation of the characteristic polynomials, hence without being necessarily conjugated over \mathbb{K} . We have to take care of two main points:

1. We want correct informations (e.g. regularity indices) before fields splittings. To do so, the parametrisations we compute are regular (they do not contain any zero divisors).
2. We want to recover usual system of RPEs after fields splittings.

In particular, the computed parametrisation will fit the following definition:

Definition 18. Let $F \in \mathbb{K}[X, Y]$ be a separable polynomial over a field \mathbb{K} . A system of *D5 rational Puiseux expansions* over \mathbb{K} of F above 0 is a set $\{R_i\}_{1 \leq i \leq \rho}$ such that:

- $R_i \in \mathbb{K}_{P_i}((T))^2$ for some square-free polynomial P_i ,
- Denoting $P_i = \prod_j P_{ij}$ the univariate factorisation of P_i over \mathbb{K} and $\{R_{ij}\}_j$ the splitting of R_i according to the decomposition $\mathbb{K}_{P_i} = \bigoplus_j \mathbb{K}_{P_{ij}}$, then the set $\{R_{ij}\}_{i,j}$ is a system of \mathbb{K} -RPE of F above 0 (as in Definition 2).

In order to deal with *all* critical points in Section 6, we will not compute the RPE's of F above 0 but above a root of a square-free factor Q of the resultant R_F . This leads us to the following definition:

Definition 19. Let $F \in \mathbb{K}_Q[X, Y]$ separable for some $Q \in \mathbb{K}[X]$ square-free. We say that F admits a system of D5-RPE's over \mathbb{K}_Q above 0 if there exists parametrisations as in Definition 18 that are regular over \mathbb{K}_Q .

Now, a system of *D5 rational Puiseux expansions* over \mathbb{K} of F above the roots of Q is a set $\{Q_i, \mathcal{R}_i\}_i$ such that:

- $Q = \prod_i Q_i$,
- \mathcal{R}_i is a system of D5 RPE's over \mathbb{K}_{Q_i} of $F(X + Z, Y) \bmod Q_i(Z)$ above 0 (in the sense of definition above).

5.2 Complexity of dynamic evaluation.

Results of [10]. We start by recalling the main results of [10], providing them only with the \mathcal{O} notation (i.e. forgetting logarithm factors). In particular, we will take $M(d) \in \mathcal{O}(d)$ in the following. In our paper, we also assume the number of variables defining triangular sets to be constant (we usually have $s = 2$ in our context).

Definition 20. An *arithmetic time* is a function $I \mapsto A_s(I)$ with real positive values and defined over all triangular sets in $\mathbb{K}[Z_1, \dots, Z_s]$ such that the following conditions hold:

1. For every triangular decomposition $I = I_1 \cap \dots \cap I_h$, we have $A_s(I_1) + \dots + A_s(I_h) \leq A_s(I)$.
2. Any addition or multiplication in \mathbb{K}_I can be made in $A_s(I)$ operations over \mathbb{K} .
3. Given a triangular decomposition $I = I_1 \cap \dots \cap I_h$, one can compute a non-critical triangular decomposition of I that refines it in less than $A_s(I)$ arithmetic operations. We denote `removeCriticalPairs` such an algorithm.
4. Given $\alpha \in \mathbb{K}_I$ and a non-critical triangular decomposition $I = I_1 \cap \dots \cap I_h$, one can compute the splitting of α in less than $A_s(I)$ operations in \mathbb{K} . We denote `Split` such an algorithm.

Theorem 7. Let $I = (P_1, \dots, P_s)$ be a triangular set, and denote $d_i = \deg_{Z_i}(P_i)$. Assuming s to be constant, one can take $A_s(I) \in \mathcal{O}(d_1 \dots d_s)$

Proof. This is a special case of the main result of [10], namely Theorem 8.1 therein. □

Proposition 13. Let $I = (P_1, \dots, P_s)$, and $a, b \in \mathbb{K}_I[Y]$ with degrees in y less than d . Assuming s constant, one can compute the extended greatest common divisor of a and b in less than $\mathcal{O}(d \cdot d_1 \dots d_s)$ operations over \mathbb{K} .

Proof. This is [10, Proposition 4.1]. □

We need a few additional results in our context. In the remaining, we focus on the case $s = 2$, denoting $I = (Q, P)$, $d_Q = \deg_{Z_1}(Q)$, $d_P = \deg_{Z_2}(P)$ and $d_I = d_Q d_P$.

Splitting all coefficients of a polynomial.

Lemma 12. There exists an algorithm `ReducePol` that, given $H \in \mathbb{K}_I[X, Y]$, returns a collection $\{(I_k, H_k)_k\}$ such that $I = \bigcap_k I_k$ is a non critical triangular decomposition and the polynomials $H_k = H \bmod I_k$ are regular over I_k . This algorithm performs at most $\mathcal{O}(d_X(H) d_Y(H) d_I)$ operations over \mathbb{K} .

Proof. The idea is similar to [10, Algorithm monic]. For each coefficient of H , we split it according to the decomposition of I found so far. For each reduced coefficient we get, we test its regularity using gcd computation. This gives us a new (possibly critical) decomposition of I . We run Algorithm `removeCriticalPairs` on it. At the end, we split H according to the found decomposition. Complexity follows from Theorem 7 and Proposition 13. □

Square-free decomposition above \mathbb{K}_I . We say that a monic polynomial $\phi \in \mathbb{K}_I[Y]$ is square-free if the ideal $I + (\phi)$ is radical. We say that $\phi = \prod_i \phi_i^{n_i}$ is the square-free factorisation of ϕ over \mathbb{K}_I if the ϕ_i are coprime square-free polynomials in $\mathbb{K}_I[Y]$ and $n_i < n_{i+1}$ for all i . We will use the following result:

Proposition 14. *Consider \mathbb{K} a field with characteristic p and $\phi \in \mathbb{K}_I[Y]$ a monic polynomial of degree d . Assuming $p = 0$ or $p > d$, there exists an algorithm **SQR-Free** that computes a set $\{(I_k, (\phi_{k,l}, M_{k,l})_l)_k\}$ such that $I = \cap_k I_k$ is a non critical triangular decomposition and $\phi_k = \prod_l \phi_{k,l}^{M_{k,l}}$ is the square-free factorisation of $\phi_k := \phi \bmod I_k$. It takes less than $\mathcal{O}(dd_I)$ operations over \mathbb{K} .*

Proof. We compute successive gcds and euclidean divisions, using Yun's algorithm [16, Algorithm 14.21, page 395] (this result is in characteristic 0, but works in positive characteristic when $p > d$). Each gcd computation is Proposition 13. We just need to add splitting steps (if needed) in between two calls. The complexity follows by using Proposition 13 in the proof of [16, Theorem 14.23, page 396], since there are less than d calls to the algorithm **removeCriticalPairs**. \square

Keeping a constant number of variables. We extend the result of Proposition 5 above \mathbb{K}_Q for some square-free polynomial Q . This requires additional attention on splittings.

Proposition 15. *Consider $I = (Q, P)$ and $\phi \in \mathbb{K}_I[Z_3]$ square-free. Denote $d = d_P \deg_{Z_3}(\phi)$, and assume \mathbb{K} to contain at least d^2 elements. There exists a Las-Vegas algorithm **BivTrigSet** that computes $(Q_k, P'_k, \psi_k)_k$ satisfying:*

- $Q = \prod_k Q_k$,
- P'_k is a squarefree polynomial of degree d over \mathbb{K}_{Q_k} ,
- ψ_k is an isomorphism between \mathbb{K}_{I_k} and $\mathbb{K}_{I'_k}$ - where $I_k = (Q_k, P \bmod Q_k, \phi \bmod Q_k)$ and $I'_k = (Q_k, P'_k)$.

It takes less than $\mathcal{O}(d^{\frac{\omega+1}{2}} d_Q)$ operations over \mathbb{K} . Given $H \in \mathbb{K}_{I_k}[X, Y]$, one can compute $\psi_k(H)$ in less than $\mathcal{O}(d_X(H) d_Y(H) d_P d d_{Q_k})$.

Proof. We follow the Las Vegas algorithm³ given in [29, Section 2.2]. First, trace computation of the monomial basis takes $\mathcal{O}(M(dd_Q))$ operations in \mathbb{K} (it is reduced to polynomial multiplication thanks to [22, Proposition 8]). Then, picking a random element A , we compute the $2d$ traces of powers of A by power projection. Methods based on [31] involve only polynomial, transposed polynomial and matrix multiplications, for a total in $\mathcal{O}(d^{\frac{\omega+1}{2}} M(dd_Q))$ operations in \mathbb{K} . Finally, our candidate for P' can be deduced via Newton's method in $\mathcal{O}(M(dd_Q))$ operations. It remains to test its square-freeness, involving gcd over \mathbb{K}_Q . It takes less than $\mathcal{O}(dd_Q)$ operations over \mathbb{K} from Proposition 13. If a factorisation of Q appears, we run some splittings and Theorem 7 concludes.

To compute ψ_k , we first need d additional traces; this is once again power projection. Then, one solves a linear system defined by a Hankel matrix (see [31, Proof of Theorem 5]). This can be done using algorithm described in [5], that reduces the problem to extended gcd computation, thus involves potential decomposition of Q . This is once again $\mathcal{O}(dd_Q)$ operations over \mathbb{K} (using **removeCriticalPairs** if needed).

To conclude, rewriting the coefficients of a polynomial $H \in \mathbb{K}_{I_k}[X, Y]$ can be done in $\mathcal{O}(d_X(H) d_Y(H) d_P d d_{Q_k})$ using e.g. Horner's scheme [26, Section 5.1.3, page 209]. \square

Remark 9. Algorithm **BivTrigSet** keeps the number of variables constant (at most two) for the triangular sets we are using during the whole algorithm. We do not work with univariate triangular sets for two reasons:

1. Computing such triangular set (starting from a bivariate one) would lead to a bound in $d_Q^{\frac{\omega+1}{2}}$, that can be $D^{\omega+1}$ when the factor Q of the resultant has high degree (see Section 6). As $\omega > 2$, this is too much.
2. Q (factor of the resultant) and P (residual extension) do not provide the same geometrical information.

Extending [19, Algorithm 1] to the D5 context.

Proposition 16. *Given $G, H \in \mathbb{K}_I[X, Y]$, one can compute a set $(I_k, G_k, H_k, U_k, V_k, \eta_k)_k$ such that:*

³here the assumption on the number of elements of \mathbb{K} is used

- $I = \cap_k I_k$ is a non critical decomposition of I ;
- $G_k = G \bmod I_k$ and $H_k = H \bmod I_k$;
- for each k , $U_k \cdot G_k + V_k \cdot H_k = X^{\eta_k} \bmod X^{\eta_k+1}$ where η_k is the least integer such that $(G_k, H_k) \cap (X^{\eta_k}) = (X^{\eta_k})$.

This can be done in less than $\mathcal{O}(d\eta)$ arithmetic operations, with $\eta = \max_k \eta_k$ and $d = \max(d_Y(G), d_Y(H))$.

Proof (sketch). As in the fast-gcd algorithm, [19, Algorithm 1] uses polynomial multiplication and euclidean division as key ingredients. Thus, the work of [10] could be adapted to this algorithm. Writting that properly would require a lot of space; as this is not a key point of this paper, we skip it. \square

5.3 Computing polygon datas in the D5 context.

We now present the main algorithm where dynamic evaluation is involved: the computation of the Newton polygon and associated datas, namely square-free decomposition of characteristic polynomials. They are grouped in algorithm **Polygon-Data** below. Given $H \in \mathbb{K}_I[X, Y]$ known with precision n , it returns a list $\{(I_i, H_i, \Delta_{ij}, \phi_{ijk}, M_{ijk})\}_k$ where:

- $I = \cap I_i$ is a non critical triangular decomposition and the non zero coefficients of $H_i := H \bmod I_i$ are invertible;
- $\{\Delta_{ij}\}_j = \mathcal{N}_n(H_i)$ is the set of edges of H_i which can be computed from $\lceil H_i \rceil^n$ - see Definition 8;
- $\phi_{\Delta_{ij}} = \prod_k \phi_{ijk}^{M_{ijk}}$ is the square-free factorisation of $\phi_{\Delta_{ij}}$ the characteristic polynomial of H_i associated to Δ_{ij} .

Algorithm: Polygon-Data(H, I, n)

Input: I a bivariate triangular set and $H \in \mathbb{K}_I[X, Y]$ known modulo X^{n+1} . We assume $n > 0$ and $d_Y(H) > 0$.

Output: A list $\{(I_i, H_i, \Delta_{ij}, \phi_{ijk}, M_{ijk})\}$ as explained above.

```

1 foreach  $(H_i, I_i)$  in ReducePol( $H, I$ ) do
2    $\{\Delta_{ij}\}_{j=1, \dots, s} \leftarrow \mathcal{N}_n(H_i)$ ; // coefficients of  $H_i$  are not zero divisors
3   for  $j = 1, \dots, s$  do
4      $\{I_i^l, \phi_{ijk}^l, M_{ijk}^l\} \leftarrow \text{SQR-Free}(\phi_{\Delta_{ij}}, I_i)$ 
5    $\{I'_m\}_m \leftarrow \text{removeCriticalPairs}(\{I_i^l\}_{i,l})$ ;
6    $\{H'_m\}_m \leftarrow \text{Split}(H_i, \{I_i^l\}_{i,l}, \{I'_m\}_m)$ ;
7   foreach  $i, j, k$  do
8      $\{\phi'_{mjk}\}_{mjk} \leftarrow \text{Split}(\phi_{ijk}^l, \{I_i^l\}_l, \{I'_m\}_m)$ ; // assuming the good subset  $\{I'_m\}_m$  is taken
9   return  $\{(I'_m, H'_m, \Delta_{i(m)j}, \phi_{mjk}, M_{i(m)jk})\}_{m,j,k}$ ; //  $i(m)$  is the correct  $i$  associated to  $m$ 

```

Proposition 17. *Algorithm Polygon-Data works correctly. It takes less than $\mathcal{O}(d_X(H) d_Y(H) d_I)$ operations in \mathbb{K} .*

Proof. Exactness and complexity of this algorithm are straightforward consequences of Proposition 14 and Theorem 7, using $\sum_{j,k,l} \deg(\phi_{ijk}^l) \leq d_Y(H)$ for any i and $\sum_{i,l} \deg(I_i^l) = \sum_m \deg(I'_m) = \sum_i \deg(I_i) = d_I$. \square

5.4 Computing half Puiseux series using dynamic evaluation.

We finally provide the D5 variant **Half-RNP3** of algorithm **Half-RNP**: here we replace univariate factorisation by square-free factorisation and use dynamic evaluation to deal with potential zero divisors.

In order to compute also the RPEs of F above the roots of any factor Q of the resultant, we are led to consider $I = (Q, P)$ instead of P as an input for **Half-RNP3**. More precisely, the input is a set F, I, n, π such that:

- $I = (Q, P)$ is a bivariate triangular set over \mathbb{K} ($P = Z_2$ for the initial call; $Q = Z_1$ is admitted);
- $n \in \mathbb{N}$ is the truncation order we will use for the powers of X during the algorithm;
- $F \in \mathbb{K}_I[X, Y]$ is separable, monic in Y , with $d_Y > 0$, and known with precision n ;

- π a pseudo-parametrisation defined by previous computations (we take $\pi = (X, Y)$ for the initial call).

The output is a set $\{I_i, \mathcal{R}_i\}_i$ such that:

- $I = \cap_i I_i$ is a non critical decomposition,
- $\mathcal{R}_i = \{R_{ij}\}$ is a set of D5-RPE's of $F_i := F \bmod I_i$ satisfying $n - v_{ij} \geq r_{ij}$ and given with precision at least $(n - v_{ij})/e_{ij} \geq r_{ij}/e_{ij} \geq 0$,

where we let $v_{ij} := v_X(\partial_Y F_i(S))$ for any Puiseux series S associated to R_{ij} .

Algorithm: Half-RNP3(F, I, n, π)

```

1 ( $I_i, H_i, \Delta_i, \phi_i, M_i$ )i ← Polygon-Data( $F, I, n$ );
2  $\{\pi_i\}_i$  ← Split( $\pi, \{I_i\}_i$ ); // taking only once each different  $I_i$ 
3 forall  $i$  do
4   if  $M_i = d_Y(H_i)$  then
5      $\pi'_i \leftarrow \lceil \pi_i(X, Y - B_i) \rceil^n$ ; //  $B_i := \text{Coeff}(H_i, Y^{d_Y(H_i)-1})/d_Y(H_i)$ 
6     if  $M_i = 1$  then  $\{I_{i11}, \mathcal{R}_{i11}\} \leftarrow \{(I_i, \pi'_i(T, 0))\}$ ;
7     else
8        $\{I_{i1k}, \mathcal{R}_{i1k}\}_k \leftarrow \text{Half-RNP3}(\lceil H_i(X, Y - B_i) \rceil^n, I_i, n, \pi'_i)$ 
9   else
10    if  $\deg(\phi_i) = 1$  then  $\xi_{i1}, I_{i1}, H_{i1}, \pi_{i1}, s_i = -\phi_i(0), I_i, H_i, \pi_i, 1$ ;
11    else // computing a primitive representation
12       $\{I_{ij}, \Psi_{ij}\}_j \leftarrow \text{BivTrigSet}(I_i, \phi_i)$ ; //  $\Psi_{ij} : \mathbb{K}_{Q_{i,j}(Z_1), P_{i,j}(Z_1, Z_2), \phi_{i,j}(Z_1, Z_2, Z)} \rightarrow \mathbb{K}_{Q_{i,j}(W_1), P'_{i,j}(W_2)}$ 
13       $\{H'_{ij}\}_j \leftarrow \text{Split}(H_i, \{I_{ij}\}_j)$ ;  $\{\pi'_{ij}\}_j \leftarrow \text{Split}(\pi_i, \{I_{ij}\}_j)$ ;
14      forall  $j$  do  $\xi_{ij}, H_{ij}, \pi_{ij} \leftarrow \Psi_{ij}(Z), \Psi_{ij}(H'_{ij}), \Psi_{ij}(\pi'_{ij})$ ;
15      forall  $j$  do //  $\Delta_i$  belongs to  $m_i a + q_i b = l_i$ ;  $u_i, v_i = \text{Bézout}(m_i, q_i)$ 
16         $\pi''_{ij} \leftarrow \pi'_{ij}(\xi_{ij}^{v_i} X^{q_i}, X^{m_i} (Y + \xi_{ij}^{u_i})) \bmod I_{ij}$ ;
17         $H''_{ij} \leftarrow \lceil H_{ij}(\xi_{ij}^{v_i} X^{q_i}, X^{m_i} (Y + \xi_{ij}^{u_i})) \rceil^{n_i} \bmod I_{ij}$ ; //  $n_i = q_i n - l_i$ 
18         $H'''_{ij} \leftarrow \text{WPT}(H''_{ij}, n_i)$ ;
19         $\{I_{ijk}, \mathcal{R}_{ijk}\}_k \leftarrow \text{Half-RNP3}(H'''_{ij}, I_{ij}, n_i, \pi''_{ij})$ ; // no recursive call if  $n_i < 0$ 
20  $\{I'_l\}_l \leftarrow \text{removeCriticalPairs}(\{I_{ijk}\}_{ijk})$ ;
21  $\mathcal{R} \leftarrow \{\}$ ;
22 forall  $i, j, k$  do  $\mathcal{R} \leftarrow \mathcal{R} \cup \text{Split}(\mathcal{R}_{ijk}, \{I'_l\}_l)$ ; // taking the subset of  $\{I'_l\}_l$  refining  $I_{ijk}$ 
23 return  $\mathcal{R}$ ; // assuming each element of  $\mathcal{R}$  is coupled to its associated  $I'_l$ 

```

We refer to the field version Half-RNP for all notations which are not specified here.

Proposition 18. *Let $Q \in \mathbb{K}[Z]$ be square-free and $F \in \mathbb{K}_Q[X, Y]$ be monic and separable in Y . The function call $\text{Half-RNP3}(F, (Q, Z), n, (X, Y))$ returns a correct answer with at most $\mathcal{O}(d_Q n d_Y^2)$ operations over \mathbb{K} .*

Proof. Just adapt the proof of Proposition 7 to the D5 context, using Propositions 15 and 17, and Theorem 7. \square

5.5 Proof of Theorem 1.

We finally conclude the proof of Theorem 1, providing the D5 variants of algorithms MonicRNP and RNP, namely algorithms MonicRNP3 and RNP3 below.

The monic case. As in Section 4, we begin with the monic case. Therein, we assume that the Hensel algorithm is a D5 version: we first use Proposition 16, and then a loop on each output for the remaining of the Hensel procedure.

Recall the notations $R_F = \text{Res}_Y(F, F_Y)$ and $v = v_X(R_F)$. We obtain the following result:

Algorithm: MonicRNP3(F, Q, n)

Input: $Q \in \mathbb{K}[Z]$ square-free, $F \in \mathbb{K}_Q[X, Y]$ separable and monic in Y , and $n \in \mathbb{N}$ big enough.

Output: $\{(Q_i, \mathcal{R}_i)\}_i$, with $Q = \prod Q_i$ and \mathcal{R}_i a system of singular parts of D5-RPEs of $F \bmod Q_i$ above 0.

```

1  $\eta \leftarrow \min(n, 6n/d_Y)$ ;  $\mathcal{R} \leftarrow \{\}$ ;
2  $\{I_i, \mathcal{R}_i\}_i \leftarrow \text{Half-RNP3}(F, (Q, Z_2), \eta, \pi)$ ; //  $I_i = (Q_i, Z_2)$ 
3  $\{F_i\}_i \leftarrow \text{Split}(F, \{Q_i\}_i)$ ;
4 forall  $i$  do
5   Keep in  $\mathcal{R}_i$  the  $R_{ij}$  such that  $v_{ij} < \eta/3$ ; // these RPEs are known with precision  $\geq 2\eta/3$ 
6   if  $\sum_{k=1}^{\lambda_i} e_{ik} f_{ik} = d_Y$  then  $\mathcal{R} \leftarrow \mathcal{R} \cup \{Q_i, \mathcal{R}_i\}$ ; continue;
7    $G_i \leftarrow \text{NormRPE}(\mathcal{R}_i, 2\eta/3)$ ;
8    $H_i \leftarrow \text{Quo}(F_i, G_i, 2\eta/3)$ ; // no splitting since  $G_i$  is monic
9    $\{Q_{ij}, G_{ij}, H_{ij}\}_j \leftarrow \text{Hensel}(F_i, G_i, H_i, n)$ ;
10  forall  $j$  do
11     $\{(Q_{ijk}, \mathcal{R}_{ijk})\}_k \leftarrow \text{MonicRNP3}(H_{ij}, Q_{ij}, n, \pi)$ ;
12     $\{\mathcal{R}'_{ijk}\} \leftarrow \text{Split}(\mathcal{R}_i, \{Q_{ijk}\}_{j,k})$ ;
13     $\mathcal{R} \leftarrow \mathcal{R} \cup \{(Q_{ijk}, \mathcal{R}_{ijk} \cup \mathcal{R}'_{ijk})_{j,k}\}$ ;
14 return  $\mathcal{R}$ 

```

Proposition 19. Assuming that $n \geq v$ and that the trailing coefficient of R_F is not a zero divisor in \mathbb{K}_Q , a function call $\text{MonicRNP3}(F, Q, n)$ returns a correct answer in less than $\mathcal{O}(d_Q d_Y n)$ operation overs \mathbb{K} .

Proof. The assumption on the trailing coefficient of the resultant of F is needed only to ensure that the truncation bound v is enough over all factors of Q . Otherwise, this is just an adaptation of the proof of Proposition 9 to the D5 context, using Propositions 18 and 16, together with Theorem 7 once again (subroutine Quo is used only with monic polynomials, and the remaining operations do not include any division). \square

The general case. Algorithm RNP3 below computes a system of singular part (at least) of D5-RPEs of a primitive polynomial F above the roots of any square-free factor Q of its resultant R_F . We follow the same strategy as in Algorithm RNP, but we take care of triangular decompositions due to division by zero divisors. In particular, we assume that algorithm Monic is a D5 version (some divisions occur while computing F_∞). Also, inversion of the RPEs of F'_∞ can lead to some triangular decompositions. However, we do not detail these further splittings for readability.

Algorithm: RNP3(F, Q, n)

Input: $Q \in \mathbb{K}[Z_1]$ square-free, $F \in \mathbb{K}[X, Y]$ separable in Y with $d_Y > 0$, and $n \in \mathbb{N}$ big enough.

Output: A system of singular parts (at least) of D5-RPEs of F above the roots of Q .

```

1  $\mathcal{R} \leftarrow \{\}$ ;  $F' \leftarrow [F(X + Z_1, Y) \bmod Q]^n$ ; // thus  $F' \in \mathbb{K}_Q[X, Y]$ 
2  $\{Q_i, F_{i,0}, F_{i,\infty}\}_i \leftarrow \text{Monic}(F', n)$ ;
3 forall  $i$  do
4    $F'_{i,\infty} \leftarrow Y^{d_Y(F_{i,\infty})} F_{i,\infty}(X, 1/Y)$ ;
5    $\{Q_{ij}, \mathcal{R}_{ij}\}_j \leftarrow \text{MonicRNP3}(F_{i,0}, Q_i, n)$ ;
6    $\{Q'_{ik}, \mathcal{R}'_{ik}\}_k \leftarrow \text{MonicRNP3}(F'_{i,\infty}, Q_i, n)$ ;
7   forall  $k$  do
8     Inverse the second element of each  $R \in \mathcal{R}'_{ik}$ ; // split  $\{Q'_{ik}, \mathcal{R}'_{ik}\}$  if required
9    $\{Q''_{il}\}_l \leftarrow \text{removeCriticalPairs}(\{Q_{ij}\}_j \cup \{Q'_{ik}\}_k)$ ;
10  forall  $k, j$  do
11     $\mathcal{R} \leftarrow \mathcal{R} \cup \text{Split}(\mathcal{R}_{ij}, \{Q''_{il}\}_l) \cup \text{Split}(\mathcal{R}'_{ik}, \{Q''_{il}\}_l)$ ;
12 return  $\mathcal{R}$ ; // assuming elements of  $\mathcal{R}$  with the same  $Q''_{il}$  are grouped together

```

Proposition 20. *Assuming that Q is a square-free factor of R_F with multiplicity $v_Q \leq n$, a function call $\text{RNP3}(F, Q, n)$ returns the correct answer in less than $\mathcal{O}(d_Q d_Y n)$ operation overs \mathbb{K} .*

Proof. The correctness follows from Propositions 11 and 19 (the trailing coefficient of the resultant of $F_{i,0}$ and $F_{i,\infty}$ is not a zero divisor by construction). The complexity follows from Propositions 3 and 19, Theorem 7, together with the relations $d_Y(F_{i,0}) + d_Y(F_{i,\infty}) = d_Y$ and $\sum_i \deg(Q_i) = d_Q$. \square

Proof of Theorem 1. The algorithm mentioned in Theorem 1 is Algorithm RNP3, run with parameters $Q = Z_1$ and $n = v$. The computation of v can be done via [19, Algorithm 1] in the aimed bound. Note that as we consider the special case $Q = Z_1$, F has coefficients over a field and this operation does not involve any dynamic evaluation. The function call $\text{RNP3}(F, Z_1, v)$ fits into the aimed complexity thanks to Proposition 20. \square

6 Desingularisation and genus of plane curves: proof of Theorem 2 and corollaries.

It's now straightforward to compute a system of singular parts of D5 rational Puiseux expansions above all critical points. We include the RPEs of F above $X = \infty$, defined as RPEs above $X = 0$ of the reciprocal polynomial $F' := X^{d_X} F(X^{-1}, Y)$. We denote by $v_\infty := v_X(R_{F'})$; it is equal to $d_X(2d_Y - 1) - \deg(R_F)$.

Definition 21. Let $F \in \mathbb{K}[X, Y]$ be a separable polynomial over a field \mathbb{K} . A D5-desingularisation of F over \mathbb{K} is a collection $\{(\mathcal{R}_1, Q_1), \dots, (\mathcal{R}_s, Q_s), \mathcal{R}_\infty\}$ such that:

- $Q_k \in \mathbb{K}[X]$ are pairwise coprime square-free polynomials such that $R_F = \prod_{k=1}^s Q_k^{v_k}$, $v_k \in \mathbb{N}^*$;
- \mathcal{R}_k is a system of singular parts (at least) of D5-RPEs of F above the roots of Q_k ;
- \mathcal{R}_∞ is a system of singular parts (at least) of D5-RPEs of F above $X = \infty$.

Note the following points:

- we can deduce from a D5-desingularisation of F the singular part of the RPE's of F above any root of R_F ,
- we allow $v_k = v_l$ for $k \neq l$ (the above factorisation of R_F we get is not necessarily a square-free factorisation).

We obtain the following algorithm:

Algorithm: Desingularise(F)

Input: $F \in \mathbb{K}[X, Y]$ separable and primitive in Y , with $d_Y > 0$.

Output: The D5-desingularisation of F over \mathbb{K}

```

1  $\mathcal{R} \leftarrow \{\}$ ;
2 forall  $(Q, n) \in \text{SQR-Free}(R_F)$  do
3    $\mathcal{R} \leftarrow \mathcal{R} \cup \text{RNP3}(F, Q, n)$ ;
4  $n \leftarrow d_X(2d_Y - 1) - \deg(R_F)$ ;
5 if  $n > 0$  then
6    $F' \leftarrow [X^{d_X} F(X^{-1}, Y)]^n$ ;
7    $\mathcal{R} \leftarrow \mathcal{R} \cup \text{RNP3}(F', Z, n)$ 
8 return  $\mathcal{R}$ 

```

Proposition 21. *Algorithm Desingularise(F) is correct and takes less than $\mathcal{O}(d_X d_Y^2)$ operations over \mathbb{K} .*

Proof. Correctness is straightforward from Proposition 20. The computation of the resultant R_F fits in the aimed bound [16, Cor. 11.21]. The complexity is then a consequence of Proposition 20, using the classical formula $\sum_k \deg(Q_k)v_k + v_\infty = \deg(R_F) + v_\infty = d_X(2d_Y - 1)$. \square

Proof of Theorem 2. It follows immediately from Proposition 21. \square

Computing the genus of plane curves: proof of Corollary 1, 2 and 3. Let $\{(Q_k, \mathcal{R}_k)\}_k$ be a D5-desingularisation of F . Since the D5-RPEs $R_{ki} \in \mathcal{R}_k$ are regular by construction, the ramification indices of all classical Puiseux series (i.e with coefficients in $\overline{\mathbb{K}}$) determined by R_{ki} are equal. If F is irreducible over $\overline{\mathbb{K}}$, the Riemann-Hurwitz formula determines the genus g of the projective plane curve defined by F as

$$g = 1 - d_Y + \frac{1}{2} \sum_k \deg(Q_k) \sum_{i=1}^{\rho_k} f_{ki}(e_{ki} - 1),$$

where f_{ki} and e_{ki} are respectively the residual degrees and ramification indices of the RPE R_{ki} . This proves Corollary 1. Corollaries 2 and 3 follow from [25, 27], where the authors show that we can reduce F modulo a well chosen small prime within the given bit complexities.

7 Factorization in $\mathbb{K}[[X]][Y]$: proof of Theorem 3.

Our aim is to compute the irreducible analytic factors of F in $\mathbb{K}[[X]][Y]$ with precision X^N , and to do so in at most $\mathcal{O}(d_Y(v + N))$ operations over \mathbb{K} , plus the cost of one univariate factorization of degree at most d_Y . To reach that aim, we use a dichotomic generalisation of Lemma 10. Given a family $F_1, \dots, F_k \in \mathbb{K}[[X]][Y]$, we denote by

$$\kappa = \kappa(F_1, \dots, F_k) := \max_{I, J} \kappa(F_I, F_J),$$

the maximum of the lifting orders being taken over all disjoint subsets $I, J \subset \{1, \dots, k\}$, with $F_I = \prod_{i \in I} F_i$.

Proposition 22. *Let $F \in \mathbb{K}[[X]][Y]$, separable of degree d . Suppose given a modular factorisation*

$$F \equiv u F_1 \cdots F_k \pmod{X^n}, \quad n > 2\kappa \quad (1)$$

where $u \in \mathbb{K}[X]$ satisfies $u(0) \neq 0$, and where for all i , either F_i or its reciprocal polynomial F_i' is monic. Then there exists uniquely determined analytic factors F_1^*, \dots, F_k^* such that

$$F = u^* F_1^* \cdots F_k^*, \quad F_i^* \equiv F_i \pmod{X^{n-\kappa}} \text{ and } u^* \in \mathbb{K}[[X]] \text{ satisfies } u^* \equiv u \pmod{X^{n-\kappa}},$$

Moreover, starting from (1), we can compute the F_i^* up to an any precision $N \geq n - \kappa$ in $\mathcal{O}(dN)$ operations over \mathbb{K} .

Proof. Existence and unicity of the lifting follow from Lemma 10. Complexity too, using a dichotomic strategy, as done for the multifactor Hensel lifting [16, Section 15.5, page 450]. \square

Remark 10. This results improves [8, Lemma 4.1], where κ is replaced by $v/2 \geq \kappa$. Note that if $\kappa = 0$, this is the classical multifactor Hensel lifting.

We only sketch a factorisation algorithm to prove Theorem 3. We will rather compute the analytic factors of F with precision $\max(v/2, N)$.

The monic case. Suppose $F \in \mathbb{K}[X, Y]$ to be separable and monic in Y . The factorization algorithm follows from a straightforward modification of algorithm `MonicRNP3`:

1. Compute v in the aimed bound.
2. Run `Half-RNP3`($F, (Z_1, Z_2), 6v/d_Y, (X, Y)$), getting the D5-RPEs $\{R_1, \dots, R_\lambda\}$ of F above 0 s.t. $v_i \leq 2v/Y$; they are known with precision at least $4v/d_Y$, and we have $\sum_{i=1}^\lambda e_i f_i \geq d_Y/2$. Complexity fits in our bound.
3. The R_i 's have coefficients in a product of fields \mathbb{K}_{P_i} , where $\sum \deg(P_i) = \sum f_i \leq d_Y$. Performing the univariate factorization of the P_i 's and splitting accordingly the R_i 's fit in the aimed bound. Hence, one can assume that the RPEs are defined over a field. Under such an assumption, they are in one-to-one correspondence with some irreducible analytic factors $F_1^*, \dots, F_\lambda^*$ of F .
4. Let $F_i := \text{NormRPE}(R_i, 4v/d_Y)$. We have by construction $F = F_1 \cdots F_\lambda H \pmod{X^{4v/d_Y+1}}$ for some polynomial H . The polynomials F_i can be computed in the aimed bound, so does H by euclidean division.

5. By Proposition 8, we have $\kappa(F_1, \dots, F_\lambda, H) < \max_i v_i \leq 2v/d_Y$. From Proposition 22, we can lift the factorisation of step 4, getting in less than $\mathcal{O}(d_Y v)$ operations a new factorisation

$$F = \tilde{F}_1 \cdots \tilde{F}_k \tilde{H} \pmod{X^{v+1}},$$

where all factors coincide with the corresponding analytic factors F_1^*, \dots, F_k^*, H^* of F up to precision at least $v - 2v/d_Y$. Moreover, the \tilde{F}_i are irreducible.

6. Previous equality forces \tilde{H} to be separable, and its degree is at most $d_Y(F)/2$. By applying recursively steps 2 to 5 on \tilde{H} (keeping $v = v_F$ but replacing d_Y with $d_Y(\tilde{H})$), we obtain within $\mathcal{O}(d_Y v)$ operations a factorisation

$$F = \tilde{F}_1 \cdots \tilde{F}_\rho \pmod{X^{v+1}},$$

where the polynomials \tilde{F}_i are in bijection with the irreducible analytic factors F_i^* of F .

7. Since $\kappa(\tilde{F}_I, \tilde{F}_J) \leq v_X \text{Res}_Y(\tilde{F}_I, \tilde{F}_J) \leq v/2$, Proposition 22 implies that $F_i^* = \tilde{F}_i \pmod{X^{v/2}}$ for all i .
8. If $N \leq v/2$, we are done. Otherwise, we apply again Proposition 22 to lift the previous factorisation up to precision $X^{N+v/2}$, getting the analytic factors of F with precision N , in less than $\mathcal{O}(d_Y N)$ operations.

The non monic case. If F is not monic (but separable), we adapt algorithm RNP3. First compute $(u, F_0, F_\infty) := \text{Monic}(F, v)$ in the aimed bound. F_0 and the reciprocal polynomial F'_∞ of F_∞ are both monic. Moreover, they are separable (because known with precision at least v). Applying the monic case strategy, we compute their irreducible factors with precision v in the aimed complexity bound. Considering the reciprocal factors of F'_∞ , we obtain a factorisation $F = u\tilde{F}_1 \cdots \tilde{F}_\rho \pmod{X^{v+1}}$, where the \tilde{F}_i are one-to-one with the irreducible factors of F . If needed, we lift further this factorization to the desired precision as in step 8 above. Theorem 3 is proved. \square

References

- [1] S. Abhyankar. *Algebraic Geometry for Scientists and Engineers*, volume 35 of *Mathematical surveys and monographs*. Amer. Math. Soc., 1990.
- [2] J.-D. Bauch, E. Nart, and H. Stainsby. Complexity of the OM factorizations of polynomials over local fields. *LMS Journal of Computation and Mathematics*, 16:139–171, 2013.
- [3] D. Bini and V. Y. Pan. *Polynomial and Matrix Computations*, volume 1 of *Progress in Theoretical Computer Science*. Birkhäuser, Saarbrücken, 1994.
- [4] A. Bostan, F. Chyzak, B. Salvy, G. Lecerf, and E. Schost. Differential equations for algebraic functions. In *Proceedings of the 2007 International Symposium on Symbolic and Algebraic Computation*, ISSAC '07, pages 25–32, 2007.
- [5] R. P. Brent, F. G. Gustavson, and D. Y. Yun. Fast solution of toeplitz systems of equations and computation of padé approximants. *Journal of Algorithms*, 1(3):259 – 295, 1980.
- [6] E. Brieskorn and H. Knörrer. *Plane Algebraic Curves*. Birkhäuser, 1986.
- [7] D. Cantor and E. Kaltofen. On fast multiplication of polynomials over arbitrary algebras. *Acta Informatica*, 28(7):693–701, 1990.
- [8] J. Cassel. *Local Fields*, volume 3 of *Student Texts*. LMS, 1986.
- [9] C. Chevalley. *Introduction to the Theory of Algebraic Functions of One Variable*, volume 6 of *Mathematical Surveys*. AMS, 1951.
- [10] X. Dahan, E. Schost, M. M. Maza, W. Wu, and Y. Xie. On the complexity of the D5 principle. *SIGSAM Bull.*, 39(3):97–98, 2005.
- [11] J. Della Dora, C. Dicrescenzo, and D. Duval. About a new method for computing in algebraic number fields. In *EUROCAL 85*. Springer-Verlag LNCS 204, 1985.

- [12] D. Duval. Diverses questions relatives au calcul formel avec des nombres algébriques. Université de Grenoble, Thèse d'État, 1987.
- [13] D. Duval. Rational Puiseux expansions. *Compositio Math.*, 70(2):119–154, 1989.
- [14] D. Duval and A. Poteaux. Death of marc rybowicz, aged 52. *ACM Commun. Comput. Algebra*, 50(4):191–191, Feb. 2017.
- [15] M. Eichler. *Introduction to the Theory of Algebraic Numbers and Functions*. Academic Press, 1966.
- [16] J. v. z. Gathen and J. Gerhard. *Modern Computer Algebra*. Cambridge University Press, New York, NY, USA, 3rd edition, 2013.
- [17] F. Kako and T. Sasaki. Solving multivariate algebraic equations by Hensel construction. *Japan J. of Industrial and Applied Math.*, 16:257–285, 1999.
- [18] H. T. Kung and J. F. Traub. All algebraic functions can be computed fast. *Journal of the Association for Computing Machinery*, 25(2):245–260, 1978.
- [19] G. Moroz and É. Schost. A fast algorithm for computing the truncated resultant. In *ISSAC '16: Proceedings of the twenty-first international symposium on Symbolic and algebraic computation*, pages 1–8, New York, NY, USA, 2016. ACM.
- [20] D. R. Musser. Multivariate polynomial factorization. *J. ACM*, 22(2):291–308, Apr. 1975.
- [21] B. M. T. O. Cormier, M. F. Singer and F. Ulmer. Linear differential operators for polynomial equations. *J. Symbolic Comput.*, 34(5):355–398, 2002.
- [22] C. Pascal and E. Schost. Change of order for bivariate triangular sets. In *ISSAC'06*, pages 277–284. ACM, 2006.
- [23] J. M. Peral. *Polígonos de newton de orden superior y aplicaciones aritméticas*. PhD thesis, Universitat de Barcelona, 1999.
- [24] A. Poteaux. *Calcul de développements de Puiseux et application au calcul de groupe de monodromie d'une courbe algébrique plane*. PhD thesis, Université de Limoges, 2008.
- [25] A. Poteaux and M. Rybowicz. Good reduction of puiseux series and complexity of the newton-puiseux algorithm over finite fields. In *ISSAC '08: Proceedings of the twenty-first international symposium on Symbolic and algebraic computation*, pages 239–246, New York, NY, USA, 2008. ACM.
- [26] A. Poteaux and M. Rybowicz. Complexity bounds for the rational newton-puiseux algorithm over finite fields. *Applicable Algebra in Engineering, Communication and Computing*, 22:187–217, 2011. 10.1007/s00200-011-0144-6.
- [27] A. Poteaux and M. Rybowicz. Good reduction of puiseux series and applications. *Journal of Symbolic Computation*, 47(1):32 – 63, 2012.
- [28] A. Poteaux and M. Rybowicz. Improving complexity bounds for the computation of puiseux series over finite fields. In *Proceedings of the 2015 ACM on International Symposium on Symbolic and Algebraic Computation*, ISSAC '15, pages 299–306, New York, NY, USA, 2015. ACM.
- [29] A. Poteaux and E. Schost. On the complexity of computing with zero-dimensional triangular sets. *Journal of Symbolic Computation*, 50(0):110 – 138, 2013.
- [30] A. Schönage and V. Strassen. Schnelle multiplikation großer zahlen. *Computing* 7, pages 281–292, 1971.
- [31] V. Shoup. Fast construction of irreducible polynomials over finite fields. *Journal of Symbolic Computation*, 17:371–391, 1994.
- [32] R. J. Walker. *Algebraic Curves*. Springer-Verlag, 1950.
- [33] M. Weimann. Bivariate factorization using a critical fiber. *Journal of Foundations of Computational Mathematics*, pages 1–45, 2016.